Efficient and scalable aggregation for large-scale data-intensive applications

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Efficient and Scalable Aggregation for Data-Intensive Applications

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Abstract

The past decade has witnessed the era of *big-data*. Due to the vast volume of data, traditional databases and data warehouses are facing problems of scalability and efficiency. As a consequence, modern data management systems that can scale to thousands of nodes, such as Apache Hadoop and Apache Spark, have emerged and quickly become the *de-facto* platforms to process data at massive scales.

In these large-scale systems, many data processing optimizations that were well studied in the database domain have now become futile because of the novel architectures and distinct programming models. In this context, this dissertation pledged to optimize one of the most predominant operations in data processing: *data aggregation* for such large-scale data-intensive systems.

Our main contributions were the *logical* and *physical optimizations* for large-scale data aggregation, including several algorithms and techniques. These optimizations are so intimately related that without one or the other, the data aggregation optimization problem would not be solved entirely. Moreover, we integrated these optimizations as essential components in our *multi-query optimization engine*, which is totally transparent to users. The engine, the logical and the physical optimizations proposed in this dissertation formed a complete package that is runnable and ready to answer data aggregation queries from users at massive scales. To the best of our knowledge, this dissertation is among the foremost to provide a complete and comprehensive solution to execute efficient and scalable data aggregation queries for large-scale data-intensive applications using MapReduce-like systems.

Our optimization algorithms and techniques were evaluated both theoretically and experimentally. The theoretical analyses showed the reasons why our algorithms and techniques are a lot more scalable and efficient than other state of the art works. The experimental results using a real cluster with synthetic and real datasets confirmed our analyses, showed a significant performance improvement and revealed various angles about our algorithms. Last but not least, our works are published as open source softwares for public usages and studies.

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I would like to dedicate this dissertation to my parents, my sister and my fiancée, Vy, who have loved and supported me unconditionally. I love you with all my heart.

Table of Contents

List of Figures

List of Tables

Chapter 1

Introduction

Data is arguably the most important asset to a company because it contains invaluable information. In large organizations, users share the same data management platform to manage and process their data, whether it is a relational database, a traditional data warehouse or a modern big-data system. Regardless of the underlying technology, users or data analytic applications would like to process their data *as fast as possible*, so that they can rapidly obtain insights and make critical decisions. Even more compelling, the current era of big-data, in which data has dramatically grown in terms of both *volume* and *value*, has seen datasets of petabytes or even zetabytes becoming the norm. The enormous amount of data puts an immense pressure on data management systems to achieve efficient data processing at such massive scales.

To response to this challenging but realistic demand, the two following requirements *must* be *together* to enable any efficient data processing at massive scales:

- A data management system that is capable of scaling up to the massive size of a large-scale cluster (hundreds, thousands or even more number of nodes).
- Scalable and efficient algorithms and optimization techniques that are able to run in parallel across the whole cluster.

Luckily, the first requirement has several adequate answers. Nowadays, modern large-scale data management systems such as Apache Hadoop [1] or Apache Spark [2] have proven that they are able to scale out to clusters of thousand nodes [3]. Companies already use these systems, or similar ones, to power their daily data processing like to compute web traffic, to visualize user patterns, *etc.*. For the second requirement, the answer is to find efficient and scalable algorithms to leverage the computing power of these systems. This is a profound mission as various data processing tasks requires their own algorithms and optimization techniques.

In this dissertation, we focus on one of the most predominant operations in data processing, *data aggregation*, or sometimes called *data summarization*. Users that interact with data, especially big-data, constantly feel the needs of computing aggregates to extract insights and obtain value from their data assets. Of course, humans can not be expected to parse through terabytes or petabytes of data. In fact, typically, users interact with data through *data summaries*. A data summary is obtained by grouping data on various *combinations* of dimensions (*e.g.*, by location and/or time), and by computing *aggregates* of those data (*e.g.*, *count*, *sum*, *mean*, *etc.*) on such combinations. These summaries, or data aggregates, are then used as input data for all kinds of purposes such as joining with other data, data visualization on dashboards, business intelligence decision making, data analysis, anomaly detection, *etc.*. From this perspective, we consider data aggregation as a crucial task that is performed extremely frequently. The workload and query templates of industrial benchmarks for databases justify this point. For instance, 20 of 22 queries in TPC-H [23] and 80 out of 99 queries in TPC-DS [22] are data aggregation queries. This bestows a great chance for optimizing data aggregation to achieve superior performance.

However, algorithms and optimization techniques available for data aggregation on modern large-scale systems are still in their infancy: they are inefficient and not scalable. In addition, despite the tremendous amount of work of the database community to come up with efficient ways to compute data aggregates, the parallel architectures and distinct programming models of these systems render those works incompatible. In other words, the problem of efficient data aggregation in large-scale systems lacks the instruments to answer the second aforementioned requirement. This is our comprehensive motivation, and this dissertation is an effort to fill in the current gap.

Thesis Statement: *We design and implement novel algorithms, optimization techniques and engines that, all together, provide automatic scalable and efficient data aggregation in large-scale systems for data-intensive applications.*

In the remainder of this Chapter, we highlight our key contributions and lay out this dissertation plan.

1.1 Contributions and Dissertation Plan

The central contribution of this dissertation was an automatic optimization that enables efficient and scalable data aggregation for large-scale data-intensive applications. This is achieved through two phases: logical optimization and physical optimization. The

whole optimization was contained in an optimization engine, which is a fundamental component of any data management system whose purpose is to find the execution plan of queries of the highest performance.

Users process their data by issuing *queries* using a specific language (*e.g.* Structured Query Language - SQL) to the data management platform. After parsing and validating users' queries to ensure that they are syntactically correct, the data management system sends these queries to its optimization engine. Here, for data aggregation queries, our optimization engine first performs the logical optimization using one of our cost-based optimization algorithms. Then, it proceeds to the physical optimization phase, in which we introduce a light-weight, cost-based optimization module. This module is capable of: *i)* selecting the most efficient from our families of physical techniques to actually materialize data aggregates; *ii)* balancing the workload across different nodes in a cluster to speed up performance. The output of the physical optimization phase is an optimized execution plan. Finally, the data management takes this plan and executes it on the cluster, using the selected physical technique of ours. All of these steps are done automatically and are completely transparent to users.

The rest of this Section is dedicated to an overview of our contributions.

1.1.1 Logical Optimization for Data Aggregation

Our optimization starts with the logical optimization phase. In this phase, data aggregation queries are logically modeled using a Directed Acyclic Graph (DAG). Because a DAG is just a logical representative of the problem, we can indeed re-use many available logical optimization algorithms from the database domain. However, none of the prior works can scale well to a large number of queries, which happens frequently (*e.g.* in ad-hoc data exploration), and/or a large number of attributes that are frequently revealed in modern datasets.

Our main contribution is to propose a new algorithm, Top-Down Splitting, which scales significantly better than state of the art algorithms. We show, both theoretically and experimentally, that our algorithm incurs in extremely small optimization overhead, compared to alternative algorithms, when producing optimized solutions. This means that, in practice, our algorithm can be applied at the massive scale that modern data processing tasks require, dealing with data of hundreds or thousands of attributes and executing several thousands of queries. Even more, this comes without any sacrifice: in general, our algorithm is able to find comparable, if not better, solutions than others as illustrated in our experimental evaluation.

1.1.2 Physical Optimization for Data Aggregation

The physical optimization phase is in charged of taking the solution from the logical optimization and deciding what is the best way for the data management system to carry it out. The output of this phase is the definitive execution strategy that is later on physically run on the cluster. Thus, the physical optimization depends exhaustively on the underlying architecture and programming model. With respect to computing data aggregates on a large-scale cluster (*e.g.* a Hadoop or a Spark cluster), the physical optimization consists of: *i)* picking the most efficient algorithm with the appropriate parameters; *ii)* balancing the workload across different nodes in a cluster to speed up performance.

Our first main contribution in this phase is that we systematically explore the design space of algorithms to physically compute aggregates through the lenses of a general trade-off model [49]. We use the model to derive the upper and lower bounds of the parallel degree and the communication cost that are inherently present in these large-scale cluster. As a result, we design and implement new data aggregation algorithms that match these bounds and swipe the design space we were able to define. These algorithms prove to be remarkably faster than prior works when properly tuned.

Our second main contribution is that, we design and implement a light-weight, cost-based optimization module, which is the heart of the physical optimization. The module collects data statistics and uses them to properly pick the most efficient algorithm and parameters. This selection process is done in a cost-based manner: the module predicts the indicative runtime of each algorithm and parameter, then chooses the best one. Last but not least, it also performs workload balancing across nodes in the cluster to further speed up the performance.

1.1.3 Multi-Query Optimization Engine

None of the above optimizations would work without an optimization engine. The purpose of this engine is to gather and orchestrate multiple types of query optimization using a unified query and data representations. Each data management system has their own optimization engine. There are two types of query optimization: single-query and multi-query. The single-query optimization treats each query separately and independently, while the multi-query one optimizes multiple queries together. Both type of optimization are important, and lacking one of them would results in a suboptimal performance. Actually, the problem of optimizing data aggregations includes both single-query and multi-query optimizations.

Our optimization algorithms can be implemented inside the optimization engine of current large-scale systems like Hadoop and Spark. However, such systems currently pro-

vide only single-query optimization engines. Thus, our main contribution here is to fill in the gap by designing and implementing a multi-query optimization engine for such systems. Our design is flexible and general that it is actually easy to implement many kinds of multi-query optimization, including ours.

1.1.4 Dissertation Plan

This dissertation is organized as follows. Chapter 2 presents the fundamental background necessary to fully appreciate the context of this dissertation.

Chapter 3 is dedicated to present our logical optimization algorithm. We introduce our formal definition of the problem, as well as thoroughly discuss the state of the art algorithms and their limitations. Then, we describe in details our algorithm and give a theoretical analysis on the best case and worst case scenarios. An experimental evaluation is conducted to evaluate the effectiveness our algorithm compared to other works.

Chapter 4 presents our contributions in physical optimization for data aggregation. The first part of this Chapter describes the mathematical model that we use to calculate the bounds of data aggregation algorithms and derive its design space to match these bounds. This part is ended with an experimental evaluation to demonstrate the competence of our algorithms. The second part of this Chapter is allocated for the heart of the physical optimization: the design and implementation of the cost-based optimization module. Using our data aggregation algorithms, we show on synthetic and real datasets that our design is very light-weight and has low optimization latency.

Chapter 5 describes the design and implementation of our multi-query optimization engine. We also show how to implement the data aggregation optimization in our engine, including our techniques as well as other works. The experiments show not only the flexibility and generality of our engine, but also the end-to-end evaluation of different logical optimization algorithms to validate our work in Chapter 3.

The last Chapter, Chapter 6 of this dissertation summarizes the main results we obtained. In the last part of this Chapter, we provide a set of possible future directions and discuss our intuitive idea.

Chapter 2

Background

In this Chapter, we presents the fundamental background on data aggregation and query optimization, both logical and physical. Because the physical optimization is tied with the underlying data management systems, we also cover the basis of our chosen programming models and execution engines for data-intensive applications with a discussion about reasoning behind our choices.

2.1 Data Aggregation

In a data management system, data aggregation maintains a significant portion of user queries [71]. The family of data aggregation queries consists of four operators: *Group By*, *Rollup*, *Cube* and *Grouping Sets*. A Group By operator finds all records having identical values with respect to a set of attributes, and computes aggregates over those records. For instance, consider a table *CarSale (CS)* with two attributes *model (M)*, and *package (P)*, the query `Select M, Count(*)From CarSale Group By (M)` counts the volume of car sales for each model. The Group By operator is the building block of data aggregation, as all other operators are generalizations of Group By. For this reason, the multiple data aggregation query optimization can be also called the *multiple Group By query optimization*.

A Cube operator (introduced by Gray *et al.* [64]) computes Group Bys corresponding to all possible combinations of a list of attributes. For instance, a Cube query like `Select M, P, Count(*)From CS Group By Cube(M, P)` can be rewritten into *four* Group By queries:

```
Q1: Select M,P,Count(*) From CS Group By(M, P)
Q2: Select M, Count(*) From CS Group By(M)
Q3: Select P, Count(*) From CS Group By(P)
```

```
Q4: Select Count(*) From CS Group By(*)
```

The Group By (*) in *Q4* denotes an ***all*** Group By, (or sometimes called empty Group By), in which all records belong to a single group.

A Rollup operator [64] considers a list of attributes as different levels of one dimension, and it computes aggregates along this dimension upward level by level. Thus a Rollup query like `Select M, P, Count(*)From CS Group By Rollup(M, P)` computes volume sale for Group Bys `(M, P)`, `(M)`, `(*)` (or Q1, Q2, Q4) in the above example.

The Rollup and Cube operators allow users to compactly describe a large number of combinations of Group Bys. Numerous solutions for generating the whole space of data Cube and Rollup have been proposed [26, 37, 58, 66, 74]. However, in the era of *"big data"*, datasets with hundreds or thousands of attributes are very common (*e.g.* data in biomedical, physics, astronomy, *etc.*). Due to the large number of attributes, generating the whole space of data Cube and Rollup is inefficient. Also, very often users are not interested in the set of all possible Group Bys, but only *a certain subset*. The Grouping Sets operator facilitates this preference by allowing users to specify the exact and arbitrary set of desired Group Bys. In short, Cube, Rollup and Grouping Sets are convenient ways to declare multiple Group By queries.

Example 1: Consider a scenario in medical research, in which there are records of patients with different diseases. There are many columns (attributes) associated with each patient such as *age*, *gender*, *city*, *job*, *etc.* A typical data analytic task is to measure correlations between the diseases of patients and one of the available attributes. For instance, heart attack is often found in elderly people rather than teenagers. This can be validated by obtaining a data distribution over two-column Group By *(disease, age)* and by comparing the frequency of heart attack of elderly ages ($age \geq 50$) versus teenagers age ($12 \leq age \leq 20$). Typically, for newly developed diseases, a data analyst would look into many possible correlations between these diseases and available attributes. A Grouping Sets query allows her to specify different Group Bys like *(disease, age)*, *(disease, gender)*, *(disease, job)*, *etc.*

2.2 Large-Scale Data Processing Model

In this Section, we introduce the fundamental of the MapReduce [60] programming model and its extension: the Resilient Distributed Datasets [17]. In data-intensive applications, these two models are the most popular tools to handle a vast amount of data due to its outstanding scalability and cost-effectiveness, and they are our models of choice.

2.2.1 MapReduce Model Fundamentals

MapReduce is a batch-oriented large-data processing model that was put forward by Google [60]. The MapReduce programming model is useful in a wide range of applications, for example: to extract, transform and load data (ETL); compute the Page-rank algorithm; sort terabytes or petabytes of data; distributed pattern-based search; *etc.*. Over the time, MapReduce has evolved into being the primary tool to turn *big data* of companies into useful information, *i.e.* to perform data processing.

The MapReduce programming model has two main phases: *map* and *reduce*. The map phase processes input records (in the form of $\langle \text{key}, \text{value} \rangle$ pairs) to produce a list of intermediate $\langle \text{key}, \text{value} \rangle$ pairs. Before the map phase, the input data is distributed across a cluster of machines as *input splits* [19]. A node that is assigned to handle the map phase of an input split is called *mapper*. Similarly, a node that is assigned to handle the reduce phase is called *reducer*. The mapper outputs are *partitioned* to reducers by a *partitioner* in a way such that pairs with the same intermediate *key* go to the same reducer. The transferring data from mappers to reducers is called the *shuffle* phase, and is done automatically. After collecting all $\langle \text{key}, \text{value} \rangle$ pairs from every mapper through network, a reducer sorts its input data by the intermediate key and constructs a list of values for that key. Finally, a reducer processes each key and its list of values to produce the final results. In brief, two phases transform the data as below:

$$map(K_1, V_1) -> [K_2, V_2]^1$$
$$reduce(K_2, [V_2]) -> [K_3, V_3]$$

In MapReduce model, there is an essential function called *combine* that lowers the amount of intermediate data considerably. A *combiner* is like a mini-reducer that runs in the map phase to combine intermediate data locally before sending them over network to reducers. An example is to `count` the total request of a web service for each day over a period of time. To reduce intermediate data, the combiner partially computes the number of requests for each day within the local mapper, and sends them to reducers. The reducers then can assemble the partial results to produce the total number of requests. The map phase now can be written as: $map(K_1, V_1) -> [K_2, V_2] -> combine(K_2, [V_2]) -> [K_2, V_2]$. However if the reduce function (*e.g. count unique*) cannot construct final results from partial results, the combiner cannot be used.

2.2.2 Resilient Distributed Datasets Fundamentals

MapReduce is a simple programming model for batch processing, yet it is also efficiently applicable to a wide range of applications. Nevertheless, there are still other major fields

[1] The bracket denotes a list

of applications that MapReduce struggles to run efficiently. In this context, Resilient Distributed Datasets [17] (RDDs) extends the data flow programming model introduced by MapReduce [60] to provide a common programming model to solve these diverse distributed computation problems. RDD is a fault-tolerant, parallel data structure which let users explicitly store data in memory or on disk, manage data partitions, and manipulate them through a rich set of operators. RDD can capture most current specialized models and new applications like streaming, machine learning or graph processing.

An RDD is an immutable, partitioned set of records. An RDD can only be created through a set of operations, called *transformation*, from only two sources: the storage input and other RDDs. Some transformations are: *map*, *filter*, *union*, *groupByKey*, *etc.*.

The connection of an RDD and its parent RDDs is represented by *dependencies*. There are two kinds of dependencies: narrow dependency where each partition of the parent RDD is used by at most one child RDD, wide dependency where each partition of parent RDD is used by multiple child RDDs. For example, a filter transformation creates a narrow dependency while a `groupByKey` transformation creates a wide dependency. The wide dependency is an abstraction of the *shuffle* phase in MapReduce, while the narrow dependency is an abstraction of computation in the *map* and *reduce* phase. This is the reason why RDD is an extension of the MapReduce model. Besides, these dependencies form a Directed Acyclic Graph (DAG) of RDDs.

RDDs provides users two useful features: persistence and partitioning. Persistence is similar to caching, and very helpful when an RDD is reused many times. When persisting an RDD, each node stores any partitions of that RDD that it computes in memory and reuses them in other transformations. This allows future transformations to be much faster. Persisting is a key tool for iterative algorithms and fast interactive usage. Partitioning lets users manage how their data is split and parallelized, which is necessary for some optimizations required location information, for example, join is a transformation which would perform significantly better if having information about data location.

All transformations in Spark are *lazy*, which means Spark does not compute their results right away. Instead, they just remember the transformations applied to some base datasets using the DAG. The transformations are only computed when an *action* is called. To obtain the result from processing the data, users call an action which triggers a Spark job. Table 2.1 gives us some basic transformations and actions inside Spark.

To summarize, each RDD is characterized by five main properties:

- A list of partitions.
- A function to compute each partition (each split).
- A list of dependencies on other RDDs.

Table 2.1: Some transformations and actions in Spark

Transformation
map(f : $T \Rightarrow$U) : $RDD[T] \Rightarrow$RDD[U]
filter(f : $T \Rightarrow$Bool) : $RDD[T] \Rightarrow$RDD[T]
flatMap(f : $T \Rightarrow$Seq[U]) : $RDD[T] \Rightarrow$RDD[U]
sample(fraction : Float) : $RDD[T] \Rightarrow$RDD[T]
groupByKey() : $RDD[(K, V)] \Rightarrow$RDD[(K, Seq[V])]
reduceByKey(f : $(V, V) \Rightarrow$V) : $RDD[(K, V)] \Rightarrow$RDD[(K, V)]
union() : $(RDD[T], RDD[T]) \Rightarrow$RDD[T]
join() : $(RDD[(K, V)], RDD[(K, W)]) \Rightarrow$RDD[(K, (V, W))]
cogroup() : $(RDD[(K, V)], RDD[(K, W)]) \Rightarrow$RDD[(K, (Seq[V], Seq[W]))]
crossProduct() : $(RDD[T], RDD[U]) \Rightarrow$RDD[(T, U)]
mapValues(f : $V \Rightarrow$W) : $RDD[(K, V)] \Rightarrow$RDD[(K, W)]
sort(c : Comparator[K]) : $RDD[(K, V)] \Rightarrow$RDD[(K, V)]
partitionBy(p : Partitioner[K]) : $RDD[(K, V)] \Rightarrow$RDD[(K, V)]
Actions
count() : $RDD[T] \Rightarrow$Long
collect() : $RDD[T] \Rightarrow$Seq[T]
reduce(f : $(T, T) \Rightarrow$T) : $RDD[T] \Rightarrow$T
lookup(k : K) : $RDD[(K, V)] \Rightarrow$Seq[V]
save(path : String) : Outputs RDD to a storage system, e.g., HDFS

- A partitioner to partition its data.
- A list of preferred data locations for computation.

2.2.3 Discussion of Our Choices

Nowadays, both MapReduce and RDD are the most popular programming models for large-scale data processing due to its three main advantages: high scalability, fault-tolerance and high cost-effectiveness. The computation in both phases of MapReduce is distributed to many nodes in a cluster, which helps producing a great parallelism for a single job. If a node fails, the fault-tolerance mechanism will schedule another node to re-compute its work. If an organization wants to expand their cluster, they may easily add more nodes to further speed up the job. This horizontal scalability allows clusters to scale up to several ten thousands node [3]. Moreover, nodes in a cluster can be built from commodity hardware, which is remarkably cheaper than specialized hardware like mainframes and supercomputers [60].

A MapReduce program can be easily expressed using RDDs. In addition, if we look at two RDDs that are connected by a wide dependency, this is basically similar to a MapReduce program. Again, RDD is an extension of MapReduce with other useful features, but the

fundamental difference is that: the computation of RDDs happens not necessarily in two phase like MapReduce, but in an arbitrary number of phases. Thus, RDD also inherits the advantages of MapReduce, and is quickly becoming the next most popular tool for large-scale data processing.

With that being said, since this dissertation is about efficient and scalable data aggregation, we strongly believe that building physical optimization of our work upon the two most popular models, MapReduce and RDD, would allow us to make an immense impact to both the research and industrial communities.

2.3 Execution Engines, High-level Languages and Query Optimization

2.3.1 Execution Engines

Since the time MapReduce programming model was first introduced by Google [60], there have been tons of effort to provide different implementations for it. These implementations are called *execution engines*, and they allow users to write a MapReduce program using their effortless application programming interfaces (APIs). After that, they execute the program on a cluster and automatically manage scalability and fault-tolerance. The most well-known execution engine for MapReduce is *Apache Hadoop* [1, 75], and for RDD, it is *Apache Spark* [2].

2.3.2 High-level Languages

Immediately, users find that writing a MapReduce (or RDD) program is low-level, tedious and error-prone. In addition, users have to optimize their programs themselves, which proves to be troublesome and difficult from time to time. Therefore, users really like to write their programs in form of *queries* using a high-level language (*e.g.* SQL) and have their queries automatically optimized. This allow users to think about the semantic of their programs, not about the details of the underlying system and its APIs. Unsurprisingly, these two features are long-established in traditional database management systems, and contribute enormously to their success. Thus, for MapReduce and RDDs to reach the same height of success, there have been efforts to answer these requirements, such as:

- *Apache Pig* [8] with its high-level language called *Pig Latin* [73] for MapReduce.
- *Apache Hive* [7] with its HiveQL, a variant of SQL, for MapReduce.

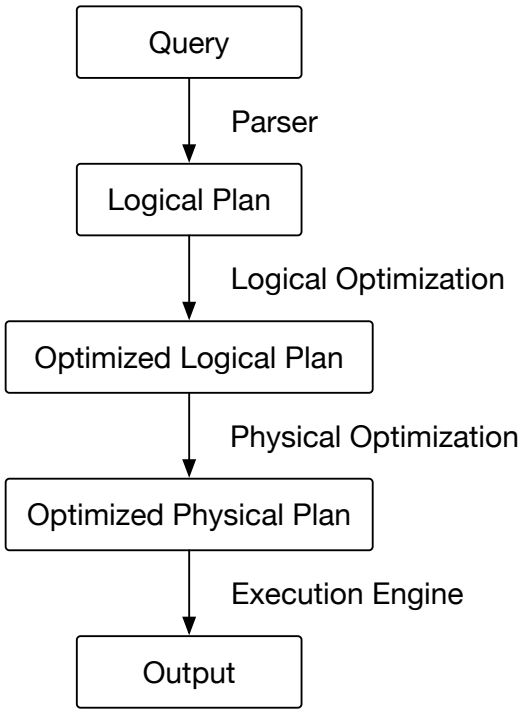

Figure 2.1: Query Processing

- *SparkSQL* [14] provides SQL engine for RDDs.

2.3.3 Query Optimization

Figure 2.1 shows the query processing pipeline. To maximize performance, queries received from users are automatically optimized through an automatic optimization engine. This engine has two phases:

- *Logical optimization*: it takes a *logical plan*, which is a logical representation of queries, and performs several optimizations over this plan. These optimizations do not care about the underlying execution engine and model but only the logical view and reasoning of these queries. For example, in any systems, running *Filter* before *Cube* would cut down a lot of execution time, since the amount of data is reduced vastly before an expensive computation is performed. Therefore, for different systems with the same logical view of queries, they can share the same logical optimization techniques.
- *Physical optimization*: optimizations in this phase focus on how to carry out the optimized logical plan, which is output of logical optimization, in an efficient way taking into account the underlying engine and model. Different models and/or

execution engines have different physical optimizations. For instance, for Hadoop MapReduce, the optimization has to decide how to form a $\langle \text{key}, \text{value} \rangle$ pair, which is a non-exist thing in traditional databases. The output of this phase is an optimized *physical plan*, and is ready for execution engine to pick up and carry out to obtain the final results.

2.4 Summary

In this Chapter, we present the basic background of data aggregation and query optimization and give principal reasons behind their enormous importance. We also shed light on the current most popular large-scale data processing tools, including the programming model as well as the execution engines. We remark that, to acknowledge users' requirements for optimizing large-scale data aggregation, this dissertation is capable of providing them:

1. Logical optimization for data aggregation (Chapter 3).
2. Physical optimization for data aggregation (Chapter 4).
3. A automatic query optimization engine that accepts high-level language queries (Chapter 5).

Chapter 3

Logical Optimization for Large-Scale Data Aggregation

In this Chapter, we present our work on the logical optimization for large-scale data aggregation.

3.1 Introduction

In this Chapter, we tackle the most general problem in optimizing data aggregation: how to efficiently compute a set of data aggregation queries. We remind that, in Section 2.1, this problem is also equivalent to the problem of how to efficiently compute a set of multiple Group By queries. This problem is known to be NP-complete ([26, 31]), and all state of the art algorithms ([24, 26, 31, 32]) use heuristic approaches to approximate the optimal solution. However, none of prior works scales well with large number of attributes, *and/or* large number of queries. Therefore, in this Chapter, we present a novel algorithm that:

- Scales well with both large numbers of attributes and numbers of Group By queries. In our experiment, the latency introduced by our query optimization algorithm is several orders of magnitude smaller than that of prior works. As the optimization latency is an overhead that we should minimize, our approach is truly desirable.
- Empirically performs better than state of the art algorithms: in many cases our algorithm finds a better execution plan, in many other cases it finds a comparable execution plan, and in only a few cases it slightly trails behind.

In the rest of the Chapter, we formally describe the problem in Section 3.2. We then discuss the related work and their limitations in Section 3.3 to motivate the need for a

new algorithm. The details of our solution with a complexity analysis are presented in Section 3.4 . We continue with our experimental evaluation in Section 3.5. A discussion about our algorithm and its extension is in Section 3.6. Finally we summarize our work and present our perspectives in Section 3.7.

3.2 Problem Statement

There are two types of query optimization: single-query optimization and multi-query optimization. As its name suggest, single-query optimization optimizes a single query by deciding, for example, which algorithm to run, configuration to use and optimized values for parameters. An example is the work in [29]: when users issue a Rollup query to compute aggregates over *day*, *month* and *year*, the optimization engine automatically picks the most suitable state of the art algorithms [28] and set the appropriate parameter to obtain the lowest query response time.

On the other hand, multi-query optimization optimizes the execution of a set of multiple queries. In large organizations, there are many users who share the same data management platform, resulting in a high probability of systems having concurrent queries to be processed. A cross industry study [71] shows that not all data is equal: in fact, some input data is *"hotter"* (*i.e.* get accessed more frequently) than others. Thus, there are high chances of users accessing these "hot" files concurrently. This is also verified by in industrial benchmarks (TPC-H and TPC-DS) in which their queries frequently access the same data. The combined outcome is that optimizing multiple queries over the *same input data* can be significantly beneficial.

The problem we address in this Chapter, the multiple Group By query optimization (MGB-QO), can come from both scenarios. From the single-query optimization perspective, any Cube, Rollup or Grouping Sets query is equal to multiple Group Bys. From the multi-query optimization perspective, the fact that many users issue one Group By over the same data means multiple Group Bys and it requires optimization. More formally, we consider an *offline* version of the problem:

- For a time window ω, without loss of generality, we assume the system receives data aggregation queries over the same input data that contains one of the following operators:
  - Group By
  - Rollup
  - Cube
  - Grouping Sets

- These queries correspond to n Group By queries $\{Q_1, Q_2, ..., Q_n\}$

In reality, hardly any query arrives at our system at the exact same time. The time window ω can be interpreted as a period of time in which queries arrive and are treated as concurrent queries. The value of ω can either be predetermined or dynamically adjusted to suit the system workload and scheduler, which lead to the *online* version of this problem. However, the online problem is not addressed in this dissertation: it remains part of our future work.

3.2.1 Definitions

We assume that the input data is a table T with m attributes (columns). Let $S = \{s_1, s_2, ..., s_n\}$ be the set of groupings that have to be computed from n queries $\{Q_1, Q_2, ..., Q_n\}$, where s_i is a subset of attributes of T. Each query Q_i is a Group By query:

Q_i: Select s_i, Count(*) From T Group By s_i

To simplify the problem, we assume that all queries perform the same aggregate measure (function) (*e.g.* Count(*)). Later in Section 3.6.2, we discuss the solution to adapt to different aggregate measures.

3.2.1.1 Search DAG

Let $Att = \{a_1, ...a_m\} = \bigcup_{i=1}^{\infty} s_i$ be the set of all attributes that appear in n Group By queries. We construct a *directed acyclic search graph* $G = (V, E)$ defined as follows. A node in G represents a grouping (or a Group By query). V is the set of all possible combinations of groupings constructed from Att plus a special node: the root node T. The root node is essentially the input data itself.

An edge $e = (u, v) \in E$ from node u to node v indicates that grouping v can be computed directly from grouping u. For instance, an edge $AB \rightarrow A$ means that grouping A can be computed from grouping AB. There are two costs associated with an edge e between two nodes: a *sort cost* $c_{sort}(e)$ and a *scan cost* $c_{scan}(e)$. If grouping AB is sorted in order of (A, B), computing grouping A would require no additional sort, but only a scan over the grouping AB. We denote this cost by $c_{scan}(e)$. However if grouping AB is not sorted, or sorted in order of (B, A), computing grouping A would require a global sort on the attribute A, incurring a sort cost $c_{sort}(e)$. The costs are of course different in two cases. We note the only exception: the root node. If input data is not sorted, then all outgoing edges from the root node have only one sort cost.

We call G a search DAG. Next, we show an example with four queries. In this example, we have an input table $T(A, B, C)$ and four Group By queries:

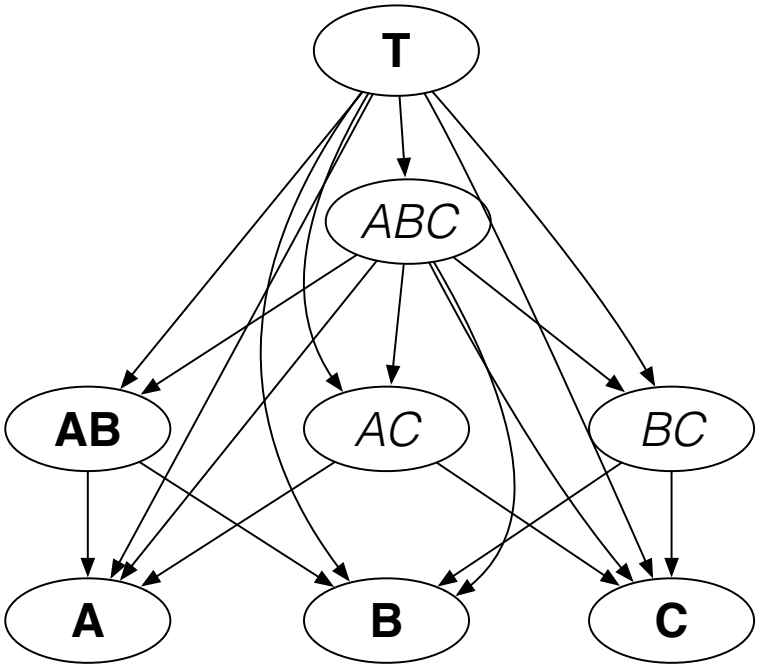

Figure 3.1: An example of a search DAG

```
Q1: Select A, Count(*) From T Group By(A)
Q2: Select B, Count(*) From T Group By(B)
Q3: Select C, Count(*) From T Group By(C)
Q4: Select A,B,Count(*) From T Group By(A,B)
```

From the above definitions, we have:

- $S = \{A, B, C, AB\}$.
- $Att = \{A, B, C\}$.
- $V = \{T, *, A, B, C, AB, AC, BC, ABC\}$.

It is easy to see that $S \subseteq V$. We call S the *terminal* (or *mandatory*) nodes: all of these nodes have to be computed and materialized as these are outputs of our Group By queries $\{Q_1, Q_2, Q_3, Q_4\}$. Other nodes in $V \setminus S$ are *additional* nodes which may be computed if it helps to speed up the execution of computing S. In this example, even though grouping ABC is not required, computing it allows $S = \{A, B, C, AB\}$ to be directly computed from ABC rather than the input table T. If the size of ABC is much smaller than T, the execution time of S is indeed reduced. Because V contains all possible combinations of groupings constructed from Att, we are sure that all possible groupings that help reduce the total execution cost are inspected. We also prune the space of V to exclude nodes that have no outgoing edges to at least one of the terminal nodes, *i.e.* these nodes

certainly cannot be used to compute S. The final search DAG for the above example is shown in Figure 3.1.

Intuitively if a grouping is used to compute two or more groupings, we want to store it in memory or disk to serve later rather than recompute it.

3.2.1.2 Problem Statement

In data management systems, the problem of multiple Group By query optimization is processed through both logical and physical optimization. In logical optimization, we set to find an optimal *solution tree* $G' = (V', E')$. The solution tree G' is a directed subtree from G, rooted at T, that covers all terminal nodes $s_i \in S$. It can be seen as a logical plan for computing multiple Group By queries efficiently. This is the main objective of this Chapter.

The physical optimization, as its name suggests, takes care of all physical details to execute the multiple Group By queries and return actual aggregates. Understandably, different data management systems have different architectures to organize their data layout, disk access, indexes, *etc.* Thus, naturally each system may have is own technique to implement the physical multiple Group By operator. For reference purpose, some example techniques are PipeSort, PipeHash [26], Partition-Cube, Memory-Cube [74] or newer technique for multiprocessors in [20] for databases, or In-Reducer Grouping [28, 80] for MapReduce and its extensions. We note that the physical optimization is not this Chapter's target: our solution is not affected by any particular physical technique. Also, regardless of the physical techniques, the grouping order is guided by the solution tree G' obtained from our logical optimization.

More formally, in the optimized solution tree G' we have:

1. $S \subseteq V' \subseteq V$.
2. $E' \subset E$ and for any edge $e(u, v) \in E'$, there is only *one* type of cost associated to edge e:

$$c(e) = \begin{cases} c_{sort}(e) \\ c_{scan}(e) \end{cases}$$

3. From any node $u \in V'$, there is at most *one* outgoing scan edge.

An optimal solution tree is the solution tree with the minimal total execution cost $C(E') = \sum_{e \in E'} c(e)$. Figure 3.2 shows an optimal solution tree for the above example. The dotted lines represent sort edges, and the solid lines show the scan edges. The bold nodes are the required grouping (*i.e.* terminal nodes). The italic node (ABC) is

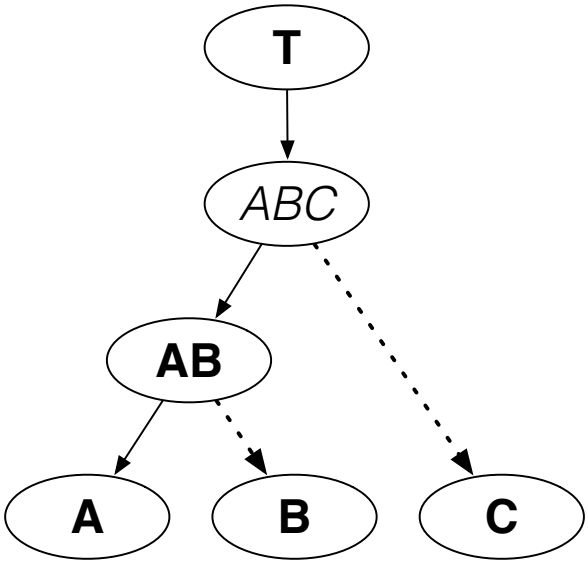

Figure 3.2: An example of a solution tree

the additional node whose computation helps to reduce the total execution cost of G'. Additional groupings BC and AC are not computed as doing so does not bring down the cost of G'.

Finding the optimal solution tree for multiple Group By queries is an NP-complete problem [31, 51]. State of the art algorithms use heuristic approaches to approximate the solution. In the next Section, we discuss in more detail why none of those algorithms scale well with large number of attributes, and/or large number of Group By queries. This motivates us to find a more scalable approach.

3.2.2 Cost model

Our primary goal is to find the solution tree G' with a small total execution cost. The total execution cost is the sum of the cost from all edges in G'. Therefore, we need a cost model that assigns the scan and sort costs to all edges in our search graph. However, our work does not depend on a specific cost model as its main purpose is to quantify the execution time of computing a node (a Group By) from another node. Any appropriate cost model for various systems like parallel databases, MapReduce systems, *etc.* can be plugged into our algorithm.

3.3 Related Work

Optimizing data aggregation in traditional databases has been one of the main tasks in database research. The multiple Group By query problem are studied through the lenses of the most general operator in data aggregation: Grouping Sets. The Grouping Sets is syntactically an easy way to specify different Group By queries at the same time, therefore all of Group By, Rollup, Cube queries can be translated directly into a Grouping Sets query.

To optimize Grouping Sets queries, the common approach is to define a directed acyclic graph (DAG) of nodes, where each node represents a single Group By appearing in the Grouping Sets query [24, 26, 31, 32, 51]. An edge from node *u* to node *v* indicates that grouping *v* can be computed from grouping *u*. For example: group BC can be computed from BCD.

There are two major differences among various works to compute Grouping Sets. The first difference is the *cost model*: how to quantify a cost (expressed as a weight of an edge) to compute a group *v* from a group *u*. PipeSort [51] sets the weight of an edge (u, v) to be proportional to the cost of re-sorting *u* to the most convenient order to compute *v*. For example, to compute BC, the main cost would be to resort ABC to BCA to compute (B, C). This is a sort cost. However, if grouping ABC is already in the sorting order of (B, C, A), the cost to compute BC would be mainly scan (hence scan cost). In contrast, [24] and [31] simplify the cost model by having only one weight for each edge (u, v), regardless of how physically v is computed: the weight of an edge (u, v) is equal to the cardinality of group *u*.

The other difference is, given a DAG of Group By nodes and weighted edges with appropriate costs, how to construct an optimal execution plan that covers all required Group Bys with the minimum total cost. This problem is proven to be NP-complete ([31, 51]), thus approximations through heuristic approaches are studied.

The work in [24] gives a simple greedy approach to address the problem. It considers Group By nodes in descending order of cardinality. Each Group By is connected to one of its super nodes. Super nodes of v is any node u such that (u, v) exists. If there are super nodes that can compute this Group By without incurring a sorting cost, it chooses the one with the least cost. This Group By becomes a scan child of its parent node. If all super nodes already have a scan child, it chooses the super nodes with the least sort cost. This approach is called *Smallest Parent*. It is simple and fast, however it does not consider any *additional* node that can help reducing the total cost. In the rest of this section, we consider algorithms that include also additional nodes.

In [26], the authors transform the problem into a *Minimal Steiner Tree* (MST) on directed graph problem. Because the cost of an edge depends on the sorting order of the parent

node, a Group By node is transformed into multiple nodes: each corresponds to a sorting order that can be generated (using permutation) from the original Group By node. Then the approach in [26] adds cost to all pairs of nodes, and uses some established approximation of MST to retrieve the optimized solution. The main drawback of this approach is that, the transformed DAG contains a huge number of added nodes and edges (because of permutation), and even a good approximation of MST problem cannot produce solutions in feasible time, as any good MST approximation is at least $\mathcal{O}(|V|^2)$ where $|V|$ is the number of nodes. For example, to compute Cube with 8 attributes, the transformed DAG consists of $109,601$ nodes and $718,178,136$ edges, w.r.t. 256 nodes and $6,561$ edges of the original DAG.

In another work [32], the authors present a greedy approach on the search DAG of Grouping Sets. Given a partial solution tree T (which initially includes only the root node), the idea is to consider all other nodes to find one that can be added to T with the most benefit. When the node x is added to T, it is first assumed to be computed from the input data set, and this incurs a sort cost. Then the algorithm in [32] tries to minimize this cost, by finding the best parent node from which x can be computed. Once x's parent is chosen, this approach finds all the nodes that are beneficial if computed from x rather than its current parent. This benefit is then subtracted by the cost of adding x to yield the total benefit of adding x to T (which the benefit value can be positive or negative). This process is repeated until it cannot find any node that brings positive benefit to add to T. The complexity of this approach is $\mathcal{O}(|V||T|^2)$ where $|V| = 2^m$, m is the number of attributes and $|T|$ is the size of the solution tree, which is typically larger than the number of terminal nodes ($|T| \geq n$). Note that while $|V|$ is much smaller than the number of nodes in [26] because of no added permutation, it is still problematic if m is large.

While the approach in [32] is more practical than the approach in [26], it cannot scale well with a search DAG of a much larger number m of attributes, in which the full space of additional nodes and edges can not be efficiently enumerated. To address this problem, [31] proposes a bottom-up approach. It first constructs a naïve plan in which all mandatory nodes are computed from the input data set. Then from all children nodes of the input data set, it considers all pairs of nodes (x, y) that can be merged. For each pair, it computes the cost of a new plan obtained by merging this pair of nodes. After that, it pick the pair, say (v_1, v_2), that has the lowest cost and replace the original plan with this new plan. In this new plan, the node $v_1 \cup v_2$ is included to the solution tree. In other words, an additional node is only considered and added if and only if it is the parent of at least two different nodes. This eliminates the task of scanning all nodes in the search DAG, making this algorithm a major improvement over previously described algorithms. At some point in time, if all the possible pairs result in a worse cost then the current plan, the algorithm stops. This algorithm calls $\mathcal{O}(n^3)$ times the procedure of

merging two nodes where n is the number of terminal nodes. The merging procedure has the complexity of $\mathcal{O}(n)$. Overall, the complexity of this algorithm is $\mathcal{O}(n^4)$.

The advantage of the algorithm described in [31] is that, it scales irrespectively to the space of $|V|$ but only to n, the number of Group By queries. If n is small, it scales better than [26,32]. However, for dense multiple Group By queries, *i.e.* large n and small m (*e.g.* computing Cube of 10 attributes results in $n = 1024$), this algorithm scale worse than [26,32]. This motivates us for a more scalable and efficient algorithm to approximate the solution tree.

3.4 The Top-Down Splitting Algorithm

In this Section, we propose a heuristic algorithm called Top-Down Splitting to find a solution tree in the multiple Group By query optimization discussed in Section 3.2. Our algorithm scales well with both large numbers of attributes and large number of Group By queries. Compared to state of the art algorithms, our algorithm runs remarkably faster without sacrificing the effectiveness. In Section 3.4.1, we present our algorithm in detail with its complexity evaluation in Section 3.4.2. Finally, we discuss the choice of appropriate values for an algorithm-wise parameter, as it affects directly the running time of our algorithm.

3.4.1 Top-Down Splitting algorithm

Our algorithm consists of two steps. The first step is to build a preliminary solution tree that consists of only terminal nodes and the root node. Taking this preliminary solution tree as its input, the second step aims to repeatedly optimize the solution tree by adding new nodes to reduce the total execution cost. While the second step sounds similar to [32], we do not consider the whole space of additional nodes. Instead, we consider only additional nodes that can evenly split a node's children into k subsets. Here k is an algorithm-wise parameter set by users. By trying to split a node's children into k subsets, we apply a structure to our solution tree: we transform the preliminary tree into a k-way tree (*i.e.* at most k fan-out). Observing the solution trees obtained from state of the art algorithms, we see that most of the times they have a relatively low fan-out k.

3.4.1.1 Constructing the preliminary solution tree

This step returns a solution tree including only terminal nodes (and of course, the root node). Later, we further optimize this solution tree. The details of this step are shown in Algorithm 1.

We sort the terminal nodes in descending order of their cardinality. As we traverse through terminal nodes in descending order, we add them to the preliminary solution tree G': we find their parent node in G' with the smallest sort cost (line 5). Obviously, nodes with smaller cardinality cannot be parents of a higher cardinality node. Thus we assure that all possible parent nodes are examined. Up to this point, we have considered only the sort cost. When all terminal nodes are added, we update the scan/sort connection between a node u and its children. Essentially, the $fix\_scan(u)$ procedure finds a child node of u that brings the biggest cost reduction when its edge is turned from *sort* to *scan* mode. The output of Algorithm 1 is a solution tree G' which is not yet optimized.

Algorithm 1 Step 1: Constructing preliminary solution

1: **function** BUILD_PRELIMINARY_SOLUTION
2: $G' \leftarrow T$
3: sort S in descending order of cardinality
4: **for** $v \in S$ **do**
5: $u_{min} = \arg\min_u c_{sort}(u, v) | u \in G'$
6: $G' \leftarrow G' \cup v$: add v to G'
7: $E' \leftarrow E' \cup e_{sort}(u_{min}, v)$
8: **end for**
9: **for** $u \in G'$ **do**
10: $fix\_scan(u)$
11: **end for**
12: **return** G'
13: **end function**

Algorithm 2 Step 2: Optimizing G'

1: **procedure** TOPDOWN_SPLIT(u, k)
2: **repeat**
3: $b \leftarrow partition\_children(u, k)$
4: **until** $(b == false)$
5: $Children = \{v_1 ... v_q\} | (u, v_i) \in E'$
6: **for** $v \in Children$ **do**
7: $topdown\_split(v, k)$
8: **end for**
9: **end procedure**

3.4.1.2 Optimizing the solution tree

In this step, we call $topdown\_split(T, k)$, with T is the root node, to further optimize the preliminary solution tree obtained in Algorithm 1. The procedure $topdown\_split(u, k)$

(Algorithm 2) repeatedly calls $partition\_children(u, k)$ (Algorithm 3) that splits the children of node u into *at most* k subsets. The function $partition\_children(u, k)$ returns ***true*** if it can find a way to optimize u, *i.e.* split children of node u into smaller subsets and reduce the total cost. Otherwise, it returns ***false*** to indicate that children of node u cannot be further optimized. We then recursively apply this splitting procedure to each child node of u. Since the flow of our algorithm is to start partitioning from the root down to the leaf nodes, we call it the *Top-Down Splitting* algorithm.

Algorithm 3 Find the best strategy to partition children of a node u to at most k subsets

```
1:  function PARTITION_CHILDREN(u, k)
2:      CN = {v_1...v_q}|(u, v_i) ∈ E'
3:      if q ≤ 1 then                          ▷ q: number of child nodes
4:          return false
5:      end if
6:      C_min = cost(G')
7:      SS ← ∅
8:      if k > q then
9:          k = q                              ▷ constraint: k ≤ q
10:     end if
11:     for k' = 1 → k do
12:         A = divide_subsets(u, k')
13:         compute the new cost C'
14:         if C' < C_min then
15:             C_min ← C'                     ▷ remember the lowest cost
16:             SS ← A                         ▷ remember new addition nodes
17:         end if
18:     end for
19:     if SS ≠ ∅ then
20:         Update G' according to SS
21:         return true
22:     else
23:         return false
24:     end if
25: end function
```

The function $partition\_children(u, k)$ (Algorithm 3) tries to split the children of u into *at most* k subsets. Each of these k subsets is represented by an additional node that is the union of all nodes in that subset. The intuition is that, instead of computing children nodes directly from u, we try to compute them from one of these k additional nodes and check if this reduces the total execution cost. Observing the solution tree obtained from state of the art algorithms, we see that in many situation, the optimal splitting strategy may not be exactly k, but a value k' ($1 \leq k' \leq k$). By trying every possible split k' from 1 to k, we compute the new total execution cost with new additional nodes, and retain the best partition scheme, *i.e.* the one with the lowest total cost. Then, we update the

solution graph accordingly: removing edges from u to children, adding new nodes and edges from u to new nodes, and from new nodes to children of u.

Algorithm 4 Dividing children into k' subsets

1: **function** DIVIDE_SUBSETS(u, k')
2: $CN = \{v_1 ... v_q\} | (u, v_i) \in E'$
3: sort CN by the descending order of cardinality
4: $C_{min} = cost(G')$
5: **for** $i = 1 \rightarrow k'$ **do**
6: $SS_i \leftarrow \emptyset$ ▷ initialize subsets i^{th}
7: **end for**
8: **for** $v \in CN$ **do**
9: $i_{min} = \arg\min_i attach(v, SS_i) | i \in 1, ...k'$
10: $SS_{i_{min}} \leftarrow SS_{i_{min}} \cup v$
11: **end for**
12: **return** $SS = \{SS_i\} \forall 1 \leq i \leq k'$
13: **end function**

The $divide\_subsets(u, k')$ (Algorithm 4) is called to divide all children of u into k' subsets and return k' new additional nodes. At first, we sort the children nodes (CN) in descending order of their cardinality. As we traverse through these children nodes, we add each child node into a subset that yields the smallest cost. The cost of adding a child node v into a subset SS_i is:

$$attach(v, SS_i) = \big[c_{sort}(u, SS_i \cup v) + c_{sort}(SS_i \cup v, v) - c_{sort}(u, v)\big]$$

Here SS_i denotes the additional node representing the i^{th} subset ($i \leq k'$). If a node v is attached to a subset SS_i, the new additional node is updated: $SS_i \leftarrow SS_i \cup v$.

Now that we have described our two steps, our algorithm is described in Algorithm 5.

Algorithm 5 Top-Down Splitting algorithm

1: $G' =$ *build_preliminary_solution()*
2: $topdown\_split(G'.getRoot(), k)$

3.4.2 Complexity of our algorithm

In this Section, we evaluate the complexity of our algorithm in the best case and the worst case scenarios. The average case complexity depends on uncontrolled factors such as:

input data distribution, relationship among multiple Group Bys, specific cost models, *etc.* We cannot compute the average complexity without making assumptions on such factors. Therefore this remains part of our future work. Empirically, we observe that in our experiments the average case leans towards the best case with just a few exceptions that are closer to the worst case.

3.4.2.1 The worst case scenario

As the first step and the second step of our algorithm are consecutive, the overall complexity is the maximum complexity of two steps. It is easy to see that the complexity of Algorithm 1 is $\mathcal{O}(n^2)$ where $n = |S|$ is the number of Group By queries.

For the second step, we first analyze the complexity of Algorithm 3: it calls $\mathcal{O}(k)$ times the $divide\_subsets$ function and it computes $\mathcal{O}(k)$ times the cost of the modified solution tree. The complexity of the $divide\_subsets$ function (*i.e.* Algorithm 4) is $\mathcal{O}(\max(k^2, kq))$. As we cannot divide q children nodes into more than q subsets, $k \leq q$. Therefore the complexity of Algorithm 4 is $\mathcal{O}(kq)$. It is not difficult to see that q is bounded by n, *i.e.* $q \leq n$. The case of $q = n$ happens when all mandatory nodes connect to the root node. Therefore the worst case complexity of Algorithm 4 is $\mathcal{O}(kn)$

Since Algorithm 3 limits itself in only modifying node u and its children, we can compute the new cost by accounting only altered nodes and edges. There are at most k new additional nodes, and there are q children nodes of node u, so computing each time a new cost of the solution tree is in $\mathcal{O}(k + q)$ time. As $k \leq q \leq n$, the complexity of computing a new cost is $\mathcal{O}(n)$, which is smaller than $\mathcal{O}(kn)$ of Algorithm 4. As such, the worst case complexity of Algorithm 3 is $\mathcal{O}(k^2 n)$

The complexity of Algorithm 2 depends on how many times $partition\_children$ is called. Let $|V'|$ be the number of nodes in the final solution tree. Clearly $topdown\_split$ is called at most $|V'|$ times, and each time it calls $partition\_children$ at least once. In order for $topdown\_split$ to terminate, $partition\_children$ has to return ***false***, and it does so in $\mathcal{O}(|V'|)$ time.

Now, for each time $topdown\_split$ is called, $partition\_children$ is called more than once if and only if it returns ***true***, which means at least an additional node is added to V'. When an additional node is added, it puts together a new subset, which consists of *at least* 2 children nodes or more. In other words, if an additional node is formed, in the worst case it applies a binary structure to the solution tree that has maximum n leaves nodes. A property of binary trees states that $|V'| \leq 2n - 1$, which means there are no more than $n-1$ additional nodes in the final solution tree. As a consequence, in the worst case, $partition\_children$ returns ***true*** in essentially $\mathcal{O}(n)$ time. Since $|V'| \leq 2n - 1$, it also returns ***false*** in $\mathcal{O}(|V'|) \equiv \mathcal{O}(n)$ time.

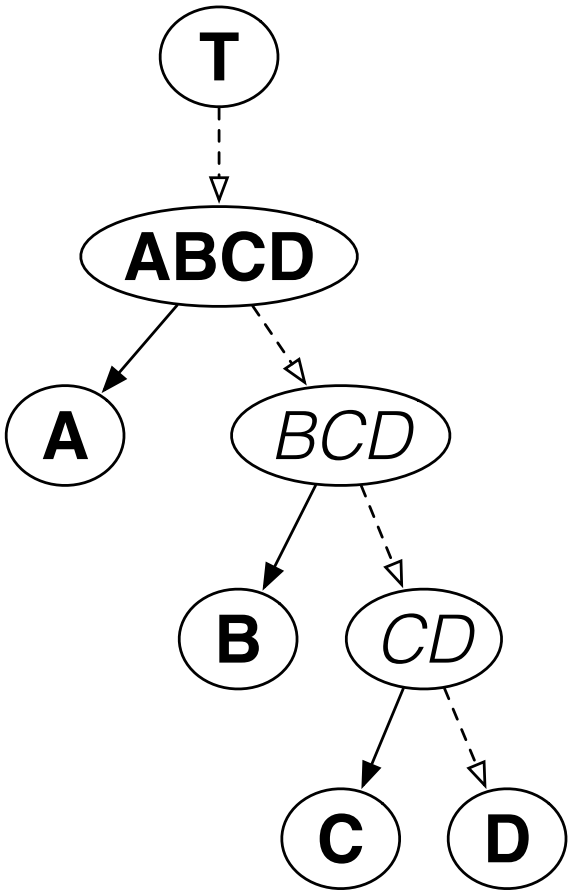

Figure 3.3: An example of worst case scenario with $k = 2$

The worst-case complexity of Algorithm 2 (*i.e.* our second step) is $\mathcal{O}(k^2n^2)$. As $\mathcal{O}(k^2n^2)$ is higher than $\mathcal{O}(n^2)$ of the first step, the worst case complexity of our algorithm is $\mathcal{O}(k^2n^2)$. Figure 3.3 shows an example of the optimized solution tree obtained in the worst case scenario.

3.4.2.2 The best case scenario

In the best case scenario, we obtain a *balanced* k-way solution tree. Figure 3.4 shows an example of such a balanced solution tree. In this scenario, Algorithm 2 calls $partition\_children$ to return ***true*** in $\mathcal{O}(log_k n)$ times instead of $\mathcal{O}(n)$ times like the worst case scenario. Therefore, the best case complexity is $\mathcal{O}(k^2 n log_k n)$.

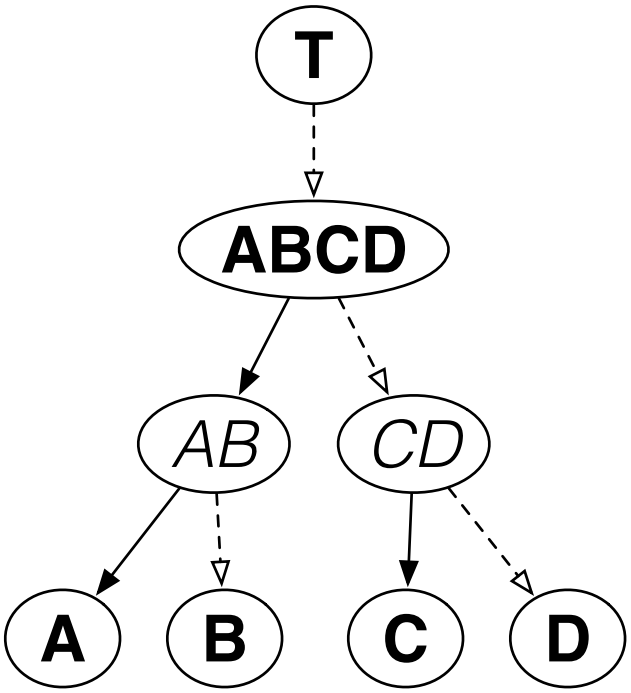

Figure 3.4: An example of best case scenario with $k = 2$

3.4.3 Choosing appropriate values of k

Our algorithm depends on an algorithm-wise parameter: k representing the fan-out of the solution tree. We observe that for solution trees obtained from state of the arts algorithm, the value of k is rather small. For example, let us consider a primary study case that motivates the work in [31]: a Grouping Sets query to compute all single-column Group By in a table with m attributes (*i.e.* the number of Group By is equal to m). In this example, small values of k such as $2 \leq k \leq 4$ are sufficient to find an optimized solution tree. In our experiments in Section 3.5, high values of k do not result in a lower cost solution tree. We note that our observation is in line to what observed in [31].

For any node u, let q_u be the number of its children. Clearly we cannot force to split u's children into more than q_u subsets, *i.e.* $k \leq q_u$. We denote $k\_max_u = q_u$. Thus, any value of k higher than q_u is wasteful, and our algorithm does not consider such values (line 9 in Algorithm 3).

On the other hand, for some node u, we cannot split its children into less than a certain number of subsets. Let us consider an example in which we want to partition 5 children nodes of $u = ABCDE$: $ABCD, ABCE, ABDE, ACDE$ and $BCDE$. Clearly splitting these children nodes into any number of subsets smaller than 5 is not possible, as merging any pairs of nodes results in the parent node u itself. In this example, 5 is the minimum number of subsets for node u. Values of k smaller than 5 result in no possible

splits. We call this the lower bound of k. To find an exact lower bound of k in a specific node u is not a trivial task. For instance, let us replace 5 children nodes of $u = ABCDE$ with: A, B, C, D, E. In this situation, the lower bound of k for node u is 2. We denote $k\_min_u = 2$

As k is an algorithm wise parameter, we have the following: $k\_min_u \leq k, \forall u \in T$. Obviously, we can set $k = max(k\_min_u)$. However, doing this is not always beneficial. Let us continue with our example, as Figure 3.3 and Figure 3.4 suggests, for node $ABCD$ and its children (A, B, C, D), $k = 2$ is sufficient to obtain an optimized solution tree; in other words, for this node, $k = 5$ is wasteful.

Algorithm 6 Adaptively dividing children

```
1:  function DIVIDE_SUBSETS(u, k')
2:      CN = {v_1...v_q}|(u, v_i) ∈ E'
3:      sort CN by the descending order of cardinality
4:      C_min = cost(G')
5:      p = k'                          ▷ p: the current number of subsets
6:      for i = 1 → p do
7:          SS_i ← ∅                    ▷ initialize subsets i^th
8:      end for
9:      for v ∈ CN do
10:         i_min = arg min_i attach(v, SS_i)|(SS_i ∪ v) ≠ u
11:         if i_min ≠ null then
12:             SS_{i_min} ← SS_{i_min} ∪ v
13:         else
14:             p ← p + 1               ▷ increase number of subsets
15:             SS_p ← v                ▷ add v to the new subset
16:         end if
17:     end for
18:     return SS
19: end function
```

In the general case, if we partition children nodes of u into a predetermined number of subsets k, *i)* for some node u it could be impossible to partition in such a way; *ii)* for some node u' it may be wasteful. Again, our observation is that most nodes have a very low fan-outs. Nodes with high upper bound of k are relatively scarce. So our strategy is to attempt partitioning children nodes of u into small numbers of subsets (*i.e.*, k is small). Whenever such a split is unachievable, we dynamically increase our number of subsets until the partition is possible. We modify Algorithm 4 to reflect the new strategy (Algorithm 6). The gist of this algorithm is that, we can attach a child node v of u to a subset SS_i if and only if $(SS_i \cup v) \neq u$. When there is no such SS_i, we add a *new* subset (*i.e.* at this node, we increase the number of subsets by 1).

3.5 Experiments and Evaluation

An optimization algorithm for the multiple Group By query problem can be evaluated from three different aspects:

- *Optimization latency*: the time (in second) that an algorithm takes to return the optimized solution tree. It is also the optimizing overhead. The lower the optimization latency, the better. This is an important metric to assess the scalability of an algorithm.
- *Solution cost*: given a cost model and a solution tree, it is the total of scan and sort cost associated to edges of the tree. A lower cost means a better tree. This metric assesses the effectiveness of an algorithm.
- *Runtime* of the solution tree: the execution time (in second) to compute n Group By queries using the optimized execution plan.

In this Section, we empirically evaluate the performance of our algorithm compared to other state of the art algorithms. The experimental results of this Section can be summarized as follows:

- The optimization latency of our algorithm is up to several orders of magnitude smaller than other algorithms when scaling to *both* large number of attributes *and* large number of Group By queries. In our experiments, the empirical results suggest that, on average, our algorithm leans towards the best case scenario more than to the worst case scenario (analyzed in Section 3.4.2).
- We do not sacrifice the effectiveness of finding an optimized solution cost for low latency. In fact, compared to other algorithms, in many cases our algorithm finds better solutions, in many other cases it finds comparable ones, and in only a few cases it slightly trails behind.
- Using PipeSort as the physical implementation to compute multiple Group By queries, we show that our algorithm can reduce the execution runtime significantly (up to 34%) compared to the naïve solution tree, in which all Group Bys are computed from the input data.

3.5.1 Experiment Setup

The experiments are run on a machine with 8GB RAM. To evaluate the latency and the solution cost of various algorithms, we synthetically generate *query templates*. Each query template consists of *i)* a list of Group By queries; *ii)* cardinalities of nodes. The cardinalities of nodes can be obtained from available datasets using the techniques described in [25, 42], or can be randomly generated (with an uniform distribution, or a

power law distribution to represent skewed data). In some situations (*e.g.* large number of attributes), we cannot effectively generate all node cardinalities. In this situation, we take the product of cardinality of each attribute in a node to be the cardinality of that node.

To evaluate the improvement that an optimized solution tree brings compared to a naïve one, we issue a query that contains all two-attribute Group Bys from the *lineitem* table of TPC-H [23]. This table contains 10 million records with 16 attributes. For each algorithm, we report its optimization latency as well as the query runtime obtained when we execute the PipeSort operator guided by its solution tree.

The state of the art algorithms that we compare are the ones presented in [32] and in [31]. We omit the algorithm presented in [26] because it is shown to be inferior to the algorithm in [32]. For convenience, we name our algorithm *Top-Down Splitting* or *TDS*, the one in [31] *Bottom-Up Merge* (*BUM*), and finally the one in [32] *Lattice Partial Cube* or *LPC*. For the case of our algorithm, as k can have multiple values, we evaluate both cases: small value $k = 3$ and large value $k = n$.

3.5.2 Cost model in our experiments

As discussed in Section 3.2.2, we need a cost model to assign scan and sort cost to edges of the search graph. In our experiments, we use a simple cost model as a representative in evaluating all different algorithms to find G'. Again, we stress that any appropriate cost model can be used. We define the two costs of an edge as follows:

- Scan cost: $c_{scan}(u \rightarrow v) = |u|$ where $|u|$ is the size (cardinality) of node u
- Sort cost: $c_{sort}(u \rightarrow v) = |u| * log_2|u|$

We assume that the cardinality of any node u is readily available: estimating $|u|$ is not the focus of our work, and we rely on works of cardinality estimation such as [25, 42]. Clearly, a bad cardinality estimation worsen the quality of a solution tree, but all algorithms suffer from the same issue.

3.5.3 Scaling with the number of attributes

In this experiment, we generate the query templates as follows: *i)* each query template consists of all single-column (or single-attribute) Group By be generated from a table T; *ii)* the number of attributes, m, in table T is from 5 to 49; *iii)* for each number of attributes, we randomly generate 100 different sets of grouping cardinalities, in which 50 sets have uniform distribution, and 50 sets have power law distribution with $\alpha = 2.5$.

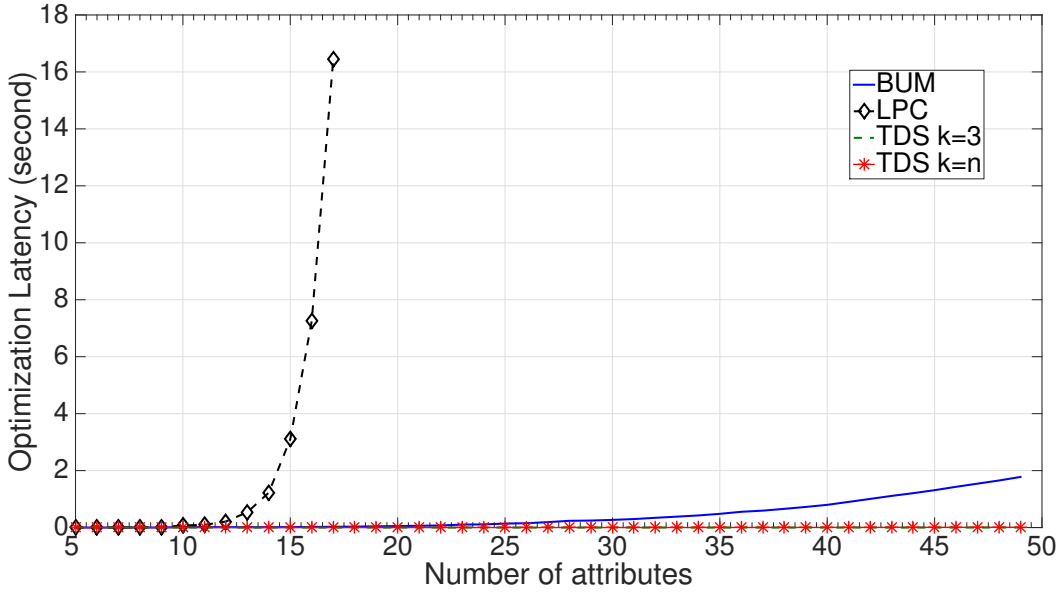

Figure 3.5: Single-attribute Group By - Optimization latency.

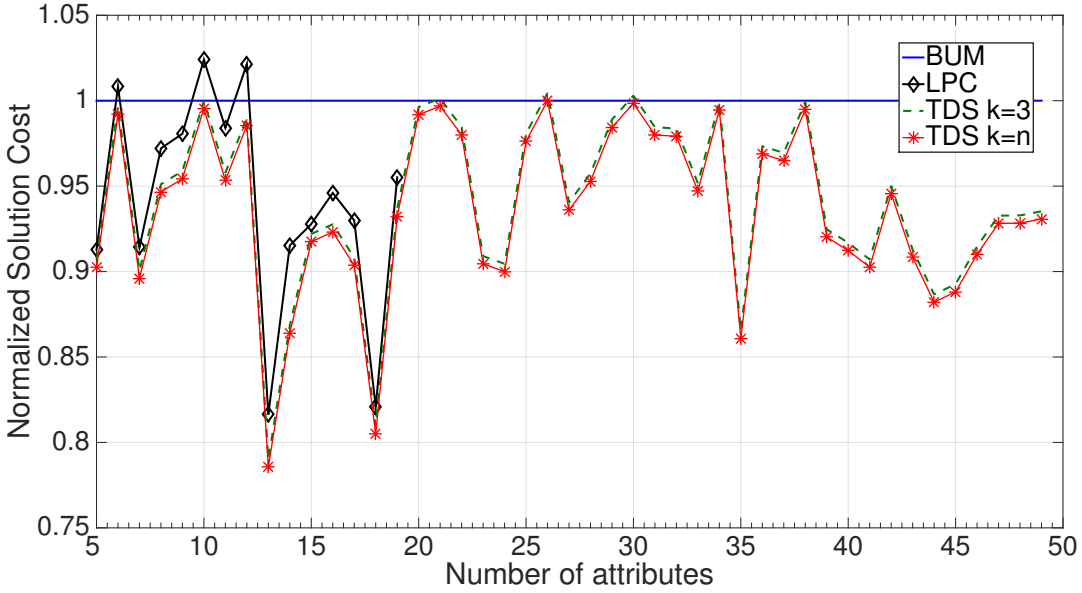

Figure 3.6: Single-attribute Group By - Normalized solution cost.

For each number of attributes and algorithm, we compute the average of the solution cost of the optimized solution tree; as well as the average of its optimization latency.

The results for optimization latency are plotted in Figure 3.5. The latency of the Lattice Partial Cube algorithm exponentially increases with the number of attributes. This is in line with its complexity of $\mathcal{O}(2^m|T|^2)$ where m is the number of attributes. For the sake of readability, we omit the latency of LPC for large number of attributes. In this

experiment, with its complexity of $\mathcal{O}(n^4)$ the Bottom-Up Merge latency scales better than LPC. Nevertheless, as the line of BUM starts to take off at the end, we expect that for large n (*e.g.* $n \geq 100$), BUM has a rather high optimization latency. Our algorithm achieves the best scalability: it takes less than 0.01 second to optimize $n = 49$ queries for both cases of k: $k = 3$ and $k = n$. Understandably, the latency of TDS with $k = n$ is higher than TDS with $k = 3$. However, as both cases have very small latencies, this is indistinguishable in Figure 3.5.

We note that the solution cost depends on the cost model, and our cost model (see Section 3.5.2) depends on grouping cardinalities, which are different in every query template we generate. Thus, we normalize every cost to a fraction of the solution cost obtained by a baseline algorithm (here we choose BUM) so that it is easier to compare the solution costs obtained by different algorithms. The normalized solution costs are shown in Figure 3.6. Due to the large amount of time for LPC to complete with large number of attributes m, we skip running LPC for $m \geq 20$. For most number of attributes, on average, our algorithm finds better solution trees than BUM, sometimes its cost is up to 20% smaller. Only in few query templates, our solution tree's cost is a little higher (within 1.5%) than BUM. Compared to LPC, TDS produces comparable execution plans. We notice several spikes of TDS and LPC. The reason is because BUM merges 2 Group Bys at a time and tends to produce uneven subsets, especially when the number of queries is an odd number. However, for some queries, BUM merging results in even subsets. In this case, its solution trees are close to TDS and LPC - thus the spikes.

Table 3.1 shows the average of solution costs obtained by each algorithm for every query template, normalized to fractional costs of BUM. Altogether, our algorithm is a little better than LPC. Also as expected, the execution cost acquired by TDS $k = 3$ is a little higher than for $k = n$, however by not much (less than 0.5%). In summary, our results indicate that for a comparable, and often lower cost than that of prior works, our approach yields substantial savings in optimization latency and scalability.

Algorithm	BUM	LPC	TDS $k = 3$	TDS $k = n$
Normalized Cost	1	0.9419	0.9406	0.9361

Table 3.1: Average solution cost - Single-attribute queries

3.5.4 Scaling with the number of queries

In this experiment, we assess the scalability of various algorithms with respect to the number of queries. To fulfill such a goal, we limit the number of attributes to be very small ($3 \leq m \leq 9$), and generate the query templates for a Cube query, *i.e.* all possible combination of Group By queries. The node cardinalities are generated similarly to Section 3.5.3.

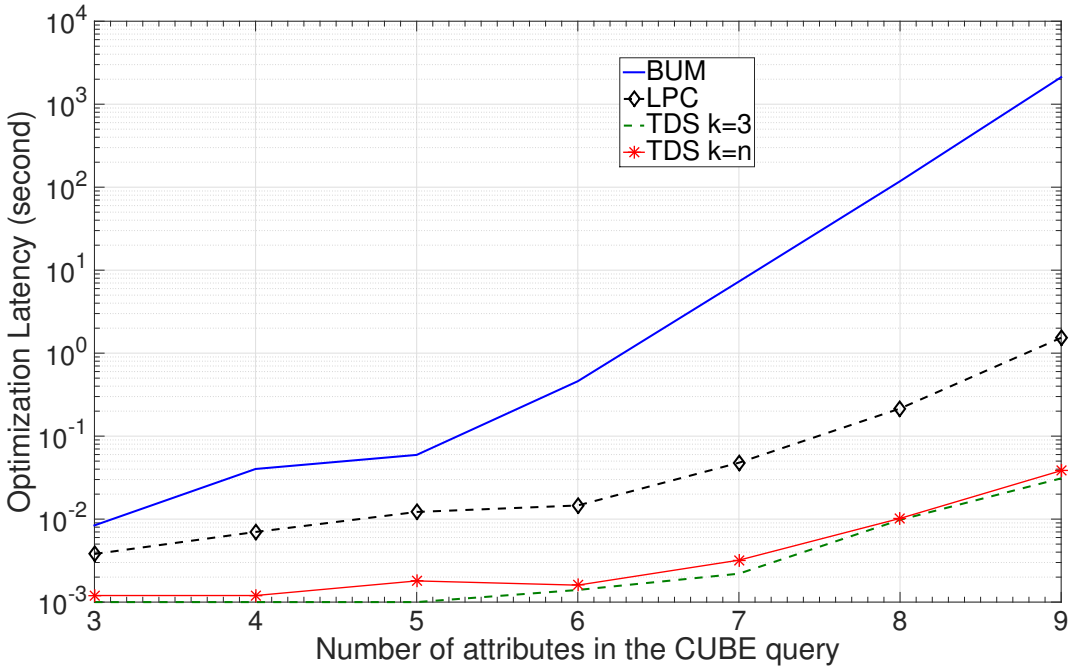

Figure 3.7: Cube queries - Optimization latency

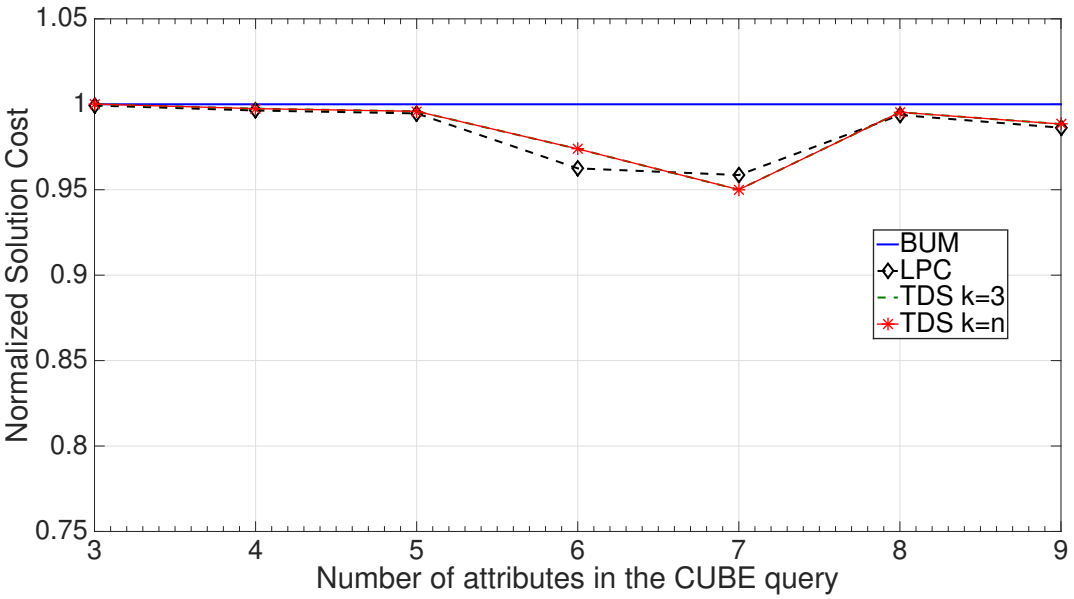

Figure 3.8: Cube queries - Normalized solution cost

The optimization latency for this experiment is in Figure 3.7. To emphasize the difference in latency between two cases TDS $k = 3$ and TDS $k = n$, we select the *log* scale for the y-axis. From Figure 3.7, we see that the latency of TDS $k = 3$ is slightly lower than TDS $k = n$. Nevertheless, in both cases our algorithm still scales remarkably better than other

algorithms. With $m = 9$, there are $2^9 = 512$ number of queries: it takes our algorithm less than 0.1 second to complete. As we mentioned in Section 3.3, in the case of dense multiple Group By queries, *i.e.* large n and small m, the BUM algorithm actually scales worse than the LPC algorithm (of which the complexity in the case of a Cube query of becomes $\mathcal{O}(n^3)$).

Figure 3.8 shows the solution cost of different algorithms in a Cube query. Despite having the highest latency and thus more time to generate optimized plans, the BUM algorithm does not produce the best solution tree (*i.e.* lowest execution cost). The reason is that BUM starts the optimization process from a naïve solution tree where all nodes are computed directly from the input data. For each step, BUM considers all possible pairs to merge and it selects the one with the lowest cost. As BUM is a gradient search approach, for large number of queries, there are too many paths that can lead to local optimum. In contrast to BUM, LPC and TDS start the optimization process from a viable solution tree T, which *i)* has much lower cost compared to the naïve solution tree; *ii)* has far less cases (*e.g.* paths) to consider. In the case of a Cube query, the initial solution tree T in LPC and TDS is closely similar to the final solution tree, with only some minor modifications. This helps both algorithms to achieve much lower latency. Between our algorithm and LPC, generally the solution tree obtained by LPC is slightly better than TDS (both cases). However, the differences are within 1.5%, which is acceptable if we want to trade effectiveness in finding solution tree for better scalability. For example, when $m = 9$, our algorithm runs in less than 0.05 second, while LPC runs in 1.5 second. Between two cases of TDS, we actually find very similar solution trees since they both start from similar preliminary trees.

3.5.5 Scaling with both number of queries and attributes

In this experiment, we compare optimization algorithms by scaling both the aforementioned factors at the same time: number of attributes and number of queries. To achieve such a goal, we design the query templates to include all two-attribute Group By queries from a table T. We set the number of attributes m from 5 to 21, and this makes the number of queries, which is $\binom{m}{2}$, increase as well. The grouping cardinalities are generated similarly to Section 3.5.3.

The optimization latency shown in Figure 3.9 exhibits the same traits that we observe in Sections 3.5.3 and 3.5.4. For low number of attributes, LPC has lower latencies compared to BUM. However, the space of additional nodes scales exponentially with the number of queries, so starting from $m = 15$ ($n = 91$), optimization latency of BUM gets smaller than that of LPC. Unsurprisingly, our algorithm has distinctly low overhead. In fact, the experimental results give us strong confidence that our algorithm is ready to scale up to hundreds or even thousands of attributes and queries. An end-to-end evaluation

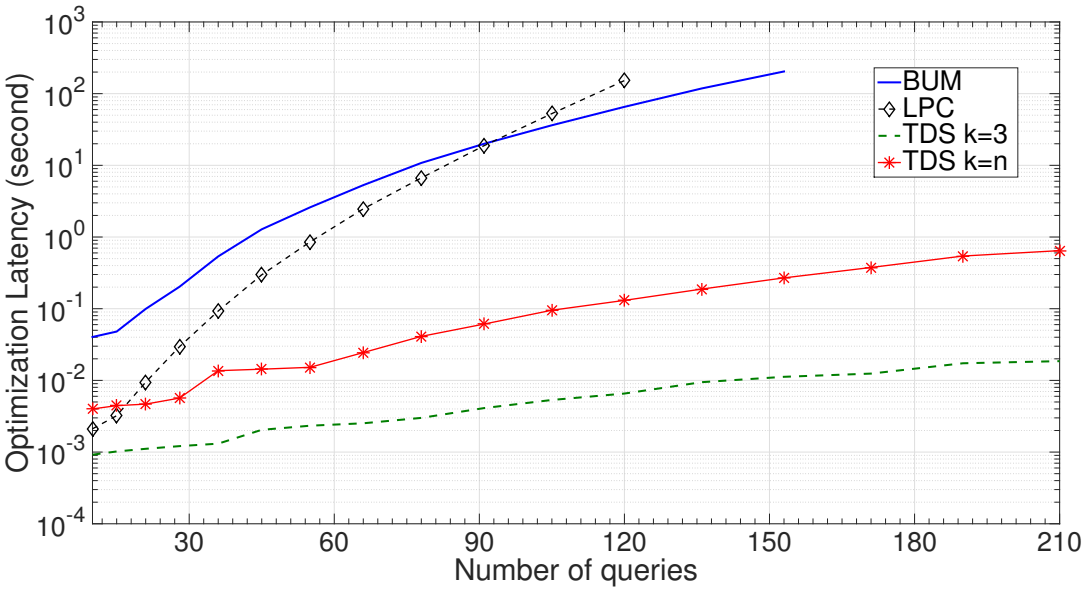

Figure 3.9: Two-attribute Group By - Optimization latency

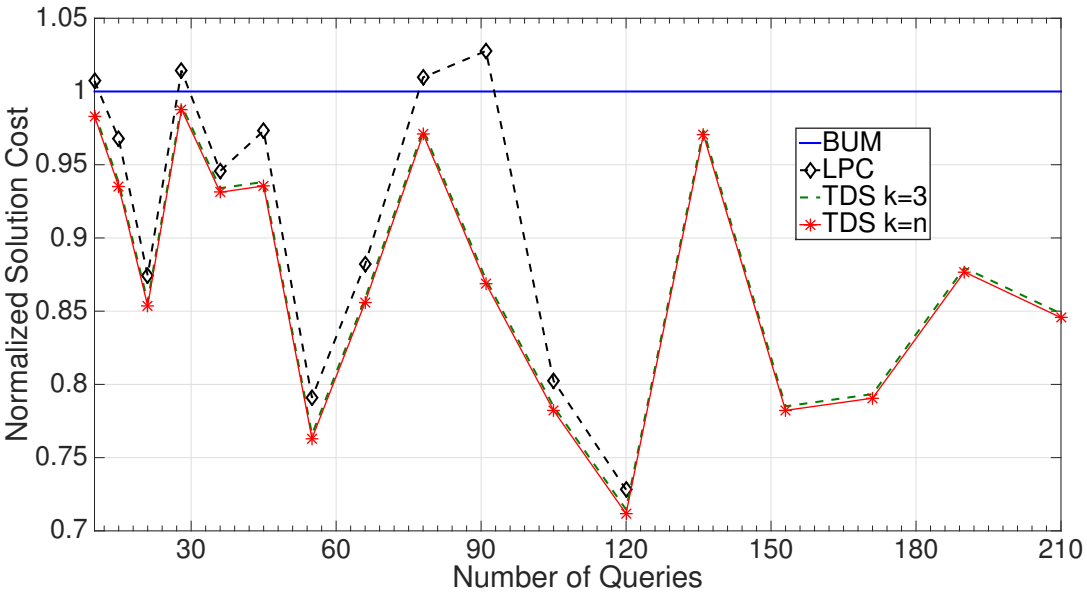

Figure 3.10: Two-attribute Group By - Normalized solution cost

of our optimization techniques handled with the physical implementation of Group By operators is part of our on-going work, which will ultimately validate the scalability and efficiency of our approach.

Figure 3.10 shows the solution cost of different algorithms. Again, we normalize it to a fraction of the BUM total cost. In some cases, we have spikes where BUM merging results in even subsets, which is also the goal of TDS. In most cases, TDS actually finds smaller

solution costs than that of BUM. Even in minor cases where TDS trails behind BUM, the difference is less than 3%, which is acceptable considering such a low optimization latency it brings. Table 3.2 presents the total average of execution cost obtained from each algorithm. An interesting observation here is that, despite having as much as 5 times the latency of TDS $k = 3$ (see Figure 3.9), the solution tree returned by TDS $k = n$ actually has less than 1% smaller total cost on average. In general, our solution trees have lower execution cost than LPC, but not by much. For some cases, both LPC and our algorithm find significantly smaller solution trees than BUM.

Algorithm	BUM	LPC	TDS $k = 3$	TDS $k = n$
Fractional Cost	1	0.9186	0.8760	0.8732

Table 3.2: Average solution cost - Two attribute queries

3.5.6 The impact of cardinality skew

As we mention in Section 3.5.3, the node cardinalities are randomly generated with two different distributions: uniform and power law with $\alpha = 2.5$. Overall, the skew introduced by the power law distribution does not affect the latency of our algorithm: on average, queries generated from both distributions have roughly the same runtime[1]. In spite of that, the total cost obtained from skewed cardinalities is generally higher than the solution cost from uniform cardinalities. On average, it is 6.8% higher, with some particular cases that are up to 18% higher. Our explanation is that, for uniform cardinalities, it is easier to *evenly* partition children nodes into subsets, while skewed cardinalities tend to return a very big subset and many small subsets. As a very big subset is used to compute other nodes, most likely it increases the solution cost by a large margin.

3.5.7 Quality of solution trees

A solution tree is a logical plan to direct the physical operator to execute multiple Group By operator. A solution tree G'_1 is of higher quality than a solution tree G'_2 if the runtime to execute G'_1 is smaller than the runtime to execute G'_2. In this experiment, we implement the PipeSort algorithm as the physical operator to execute the solution trees returned from TDS, LPC and BUM. We also execute a naïve solution tree to prepare a baseline for comparison.

The dataset we use in this experiment is the *lineitem* table from the industrial benchmark TPC-H [23]. It contains 10 million records with 16 attributes. Our query consists

[1]Uniform distribution has a slightly higher running time of 1.4%

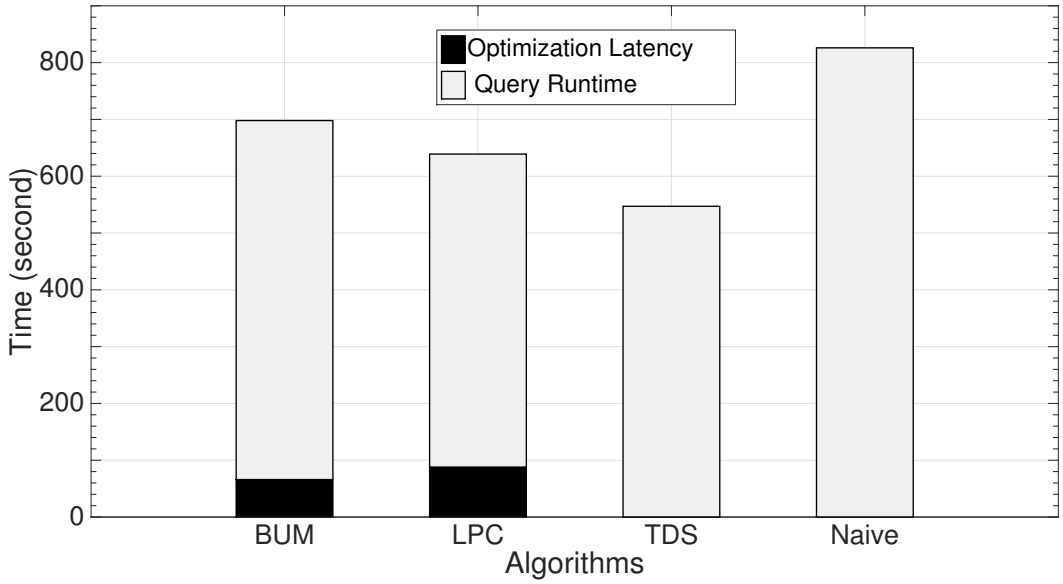

Figure 3.11: The optimization latency and query runtime.

Algorithm	BUM	LPC	TDS	Naïve
Optimization Latency (s)	66	87	0.17	0
Query Runtime (s)	632	551	547	826
Total Runtime (s)	698	638	547.17	826
Improvement (%)	15.49	22.76	33.75	0

Table 3.3: The optimization latency and query runtime

of all two-attribute Group Bys from *lineitem*. For each algorithm, we report, in Figure 3.11, its optimization latency as well as its query runtime. Detail numbers found in Table 3.3 indicate that multiple Group By query optimization techniques actually reduce the query runtime over a naïve solution. We observe that, in this workload, the optimization latency of TDS is almost 0% of the total runtime. This is in contrast to that of BUM (9.45%) and LPC (13.63%). With BUM and LPC, since the optimization latency contributes a non-negligible part to the total runtime, instead of 23.48% and 33.29% improvement respectively, they improve only 15.49% and 22.76%. Also, we note that the solution trees of LPC and our algorithm, TDS, are identical. This lead to the virtually same query runtime. Nevertheless, overall our algorithm provides greater performance boost because of its close-to-zero latency.

3.6 Discussions and Extensions

In this Section, first we discuss about the intuition of our algorithm, and attempt to explain why its solution cost may be better than other algorithms in many cases. Then we discuss about solutions to extend our algorithm to handle different aggregate functions.

3.6.1 Intuition and Discussion

In the multiple Group By query optimization problem, to design a scalable algorithm, the first building block to consider is how to explore the potential additional groupings (*i.e.* nodes) to include in a solution. As the space of additional groupings can be very large, they cannot be effectively generated. Therefore, to scale to a large number of queries and attributes, we cannot explore all possible additional groupings. A more scalable approach is to consider merging terminal nodes to form new groupings, as the subset of these additional groupings is much likely considerably smaller than their full space. Both our algorithm and BUM in [31] use this approach.

The difference between TDS and BUM is the process to construct new groupings. At each step, BUM only considers merging two groupings into a new one. Meanwhile, TDS evaluates splitting all children nodes of a grouping into at most k subsets, each with a new grouping. The implication of these steps on the algorithm's complexity is already discussed in Section 3.3 and 3.4.2. Here, we intuitively discuss why, in general, we believe that TDS can produce better solution trees than BUM. The main reason is because TDS makes a more *"global"* decision than BUM at each step of their process. When considering partitioning children nodes, TDS uses available information at the moment: *i)* cardinalities of all nodes; *ii)* associated costs to pair of nodes; *iii)* multiple ways to split. In addition, while it is making a decision of putting together a new grouping, TDS inspects the connection of this newly formed grouping with respect to all other groupings available. As a top-down approach, when TDS triggers a splitting decision in a high-level grouping (*e.g.* the root node), it dramatically decreases the total execution cost. Even though the subtree optimization might be far from optimal, early decisions are more important.

In contrast, the merging process in BUM solely depends on two individual nodes. With so little information at hand, BUM tends to trigger groupings that decrease the solution cost by a relatively small margin (because BUM is a bottom-up approach). In addition, since the initial solution tree is a naïve solution, there are so many pairs of nodes such that inspecting the potential merging of all pairs leads to a local optimum.

3.6.2 Different Aggregate Functions

The multiple Group By optimization problem is based on the premise that a Group By can be computed from the results of another Group By, instead of the input data. To assure this property, the aggregate measure (function) must be either *distributive*, or *algebraic* [64]: fortunately, almost all common aggregate functions are so. Let us consider an input data T which is split into p chunks C_i. A function F is:

- *Distributive* if there is a function G such that: $F(T) = G(F(C_1), F(C_2), ..., F(C_p))$. Usually, for many distributive functions like Min, Max, Sum, *etc.*, $G = F$. For $G \neq F$, an example is $Count$ which is also distributive with $G = sum()$.
- *Algebraic* if there are functions G and H such that: $F(T) = H(G(C_1), G(C_2), ..., G(C_p))$. Examples are $Average, MinN, MaxN, Standard\_deviation$. In function $Average$, for instance, G is to collect the sum of elements and the number of elements, while H adds up these two components and divides the global sum by the total number of elements from all the chunks to obtain final results.

Thus far, we have assumed that all Group By queries compute the same aggregate function ($Count(*)$ in our example). Typically this is the case in single-query optimization when a user issues a Grouping Sets query. Nonetheless in multi-query optimization, more often the aggregate functions are different. An easy way to adapt our solution to different aggregate functions is to separate Group By queries into groups of the same aggregate function. However, this approach may decrease the opportunity to share precomputed Group Bys, and it may end up computing a large portion of Group Bys from the input data (or from a much larger Group Bys). For instance, let us consider the following queries:

```
Q1: Select A, Count(*) From T Group By(A)
Q2: Select B, Sum(v) From T Group By(B)
```

Here v is an integer value in table T. Using the aforementioned approach we end up with both Group Bys A and B computed from the input data T.

Another approach is to apply our optimization to the set of all Group Bys queries. For a Group By, not only its aggregates are computed, but also are all those of their successors. To continue our example, our algorithm suggests computing grouping AB from T, then A and B from AB. When computing AB from T, we evaluate and store both aggregates, $Count(*)$ and $Sum(v)$. At this moment, Group By $Q1$ is obtained from grouping AB with aggregate $Count(*)$, while we use AB with aggregate $Sum(v)$ to calculate Group By $Q2$. The downside of this approach is to incur the cost of storing potentially numerous intermediate aggregates.

For systems in which the storage cost is relatively expensive compared to reading/sorting a large amount of data, the first approach may be preferred. For other systems, in which the sorting cost is relatively expensive (*e.g.* it requires a global shuffle of data over network), the second approach may be a more viable option. Facilitating users in making the right choice of approaches is the challenge that we will tackle in our future work.

3.7 Summary

Data aggregation is a crucial task to understand and interact with data. This is exacerbated by the increasingly large amount of data that is collected nowadays. Such data is often multi-dimensional, characterized by a very large number of attributes. This calls for the design of new algorithms to optimize the execution of data aggregation queries.

In this Chapter, we presented our method to address the general problem of optimizing multiple Group By queries, thus filling the gap left by current proposals that cannot scale in the number of concurrent queries or the number of attributes each query can handle. We have shown, both theoretically and experimentally, that our algorithm incurs in extremely small latencies, compared to alternative algorithms, when producing optimized query plans. This means that, in practice, our algorithm can be applied at the scale that modern data processing tasks require, dealing with data of hundreds of attributes and executing thousands of queries. In addition, our experimental evaluation illustrated the effectiveness of our algorithm to find optimized solution trees. In fact, in many cases, our algorithm outperformed others in terms of producing optimized solutions, while being remarkably faster. Finally, we discussed about the intuition behind our algorithm and the possible approaches to extend it to handle general, heterogeneous queries in terms of diversity of aggregate functions. A version of this Chapter is published in the International Conference on Data Engineering 2016 [30].

We conclude this Chapter by noting that our algorithm can be easily integrated to current optimization engines of relational databases, to traditional data warehouses or to modern big-data systems like Apache Hadoop [1], Apache Spark [2]. Instead, using our algorithm to optimize query execution on recent systems such as Hadoop, Spark and their respective high-level, declarative interfaces, requires the development of an appropriate cost model as well as an optimization engine to transform original query plans into optimized ones. We cover this in Chapter 5. In the next Chapter, Chapter 4, we continue the flow of our query optimization and go on to the next phase: physical optimization.

Relevant Publications

- D. H. Phan and P. Michiardi, "A novel, low-latency algorithm for multiple group-by query optimization", in 32nd *IEEE International Conference on Data Engineering*, 2016.

Chapter 4

Physical Optimization for Data Aggregation

This Chapter presents our contributions in physical optimization for data aggregation: *i)* our novel algorithms for data aggregation; *ii)* our light-weight, cost-based optimization module.

4.1 Design Space of MapReduce-like Rollup Data Aggregates

4.1.1 Introduction

Despite the tremendous amount of work carried out in the database community to come up with efficient ways of computing data aggregates, little work has been done to extend these lines of work to cope with massive scale. Indeed, the main focus of prior works in this domain has been on single server systems or small clusters executing a distributed database, implementing efficient implementations of Grouping Sets, Cube and Rollup operators, in line with the expectations of low-latency access to data summaries [51,58,61,65,66,74]. Only recently, the community devoted attention to solve the problem of computing data aggregates at massive scales using large-scale data-intensive computing engines such as Hadoop [1] and Spark [2] . In support of the growing interest in computing data aggregates on batch-oriented systems, several high-level languages built on top of Hadoop MapReduce, such as PIG [8] and HIVE [7], support simple implementations of, for example, the Rollup operator.

The endeavor of this work is to take a systematic approach to study the design space of the Rollup operator: besides being widely used on its own, Rollup is also a fundamental

building block used to compute more general data aggregation queries (*e.g.* Grouping Sets [30, 52]). We discuss this in more details in Section 4.3. In this Section, we study the problem of defining the design space of algorithms to implement Rollup through the lenses of a recent model of MapReduce-like systems [49]. The model explains the trade-offs that exist between the degree of parallelism that is possible to achieve and the communication costs that are inherently present when using the MapReduce programming model. In addition, we overcome current limitations of the model we use (which glosses over important aspects of MapReduce-like computations) by extending our analysis with an experimental approach. We present instances of algorithmic variants of the Rollup operator that cover several points in the design space, implement and evaluate them using a Hadoop cluster.

In summary, our contributions are the following:

- We study the design space that exists to implement Rollup and show that, while it may appear deceivingly simple, it is not a straightforward embarrassing parallel problem. We use modeling to obtain bounds on parallelism and communication costs.
- We design and implement new Rollup algorithms that can match the bounds we derived, and that swipe the design space we were able to define.
- We pinpoint the essential role of *combiners* (an optimization allowing pre-aggregation of data, which is available in real instances of the MapReduce paradigm, such as Hadoop [1]) for the practical relevance of some algorithm instances, and proceed with an experimental evaluation of several variants of Rollup implementations, both in terms of their performance (runtime) and their efficient use of cluster resources (total amount of work).
- Finally, our Rollup implementations exist in Java MapReduce and have been integrated in our experimental branch of PIG, which are available in a public repository.[1]

The remainder of this Section is organized as follows. Section 4.1.2 provides background information on the model we use in our work and presents related work. Section 4.1.3 illustrates a formal problem statement and Section 4.1.4 presents several variants of Rollup algorithms. Section 4.1.5 outlines our experimental results to evaluate the performance of the algorithms we introduce in this work. The summary of the work in this Section is presented in Section 4.4.

[1] https://bitbucket.org/bigfootproject/rollupmr

4.1.2 Background and Related Work

We assume the reader to be familiar with the MapReduce [60] paradigm and its open-source implementation Hadoop [1, 75]. First, we give a brief summary of the model introduced by Afrati *et al.* [49], which is the underlying tool we use throughout our Section. Then, we present related works that focus on the MapReduce implementation of popular data analytics algorithms.

The MapReduce model. Afrati *et al.* [49] recently studied the MapReduce programming paradigm through the lenses of an original model that elucidates the trade-off between parallelism and communication costs of single-round MapReduce jobs. The model defines the design space of a MapReduce algorithm in terms of *replication rate* and *reducer-key size*. The replication rate r is the average number of $\langle \text{key}, \text{value} \rangle$ pairs created from each input in the map phase, and represents the communication cost of a job. The reducer-key size q is the upper bound of the size of list of values associated to a reducer-key. Jobs have higher degrees of parellelism when q is smaller. For some problems, parallelism comes at the expense of larger communication costs, which may dominate the overall execution time of a job.

Afrati *et al.* show how to determine the relation between r and q. This is done by first bounding the amount of input a reducer requires to cover its outputs. Once this relation is established, a simple yet effective "recipe" can be used to relate the size of the input of a job to the replication rate and to the bounds on output covering introduced above. As a consequence, given a problem (*e.g.*, finding the Hamming distance between input strings), the model can be used to establish bounds on r and q, which in turn define the design space that instances of algorithms solving the original problem may cover.

Related work. Designing efficient MapReduce algorithms to implement a wide range of operations on data has received considerable attention recently. Due to space limitations, we cannot give justice to all works that addressed the design, analysis and implementation of graph algorithms, clustering algorithms and many other important problems: here we shall focus on algorithms to implement SQL-like operators. For example, the relational JOIN operator is not supported directly in MapReduce. Hence, attempts to implement efficient JOIN algorithms in MapReduce have flourished in the literature: Blanas *et al.* [59] studied *Repartition Join*, *Broadcast Join*, and *Semi-Join*. More recent work tackle more general cases like theta-joins [72] and multi-way-joins [50].

With respect to OLAP data analysis tasks such as Cube and Rollup, efficient MapReduce algorithms have only lately received some attention. A first approach to study Cube and Rollup aggregates has been proposed by Nandi *et al.* [69]; this algorithm, called "naive" by the authors, is called *Vanilla* in this work. MR-Cube [69] mainly focuses on algebraic

aggregation functions, and deals with data skew; it implements the Cube operator by breaking the algorithm in three phas-es. A first job samples the input data to recognize possible reducer-unfriendly regions; a second job breaks those regions into sub-regions, and generates corresponding $\langle \text{key}, \text{value} \rangle$ pairs to all regions, to perform partial data aggregation. Finally, a last job reconstructs all sub-regions results to form the complete output. The MR-Cube operator naturally implements Rollup aggregates. However in the special case of Rollup, the approach has two major drawbacks: it replicates records in the map phase as in the naive approach and it performs redundant computation in the reduce phase.

For the sake of completeness, we note that one key idea of our work (in-reducer grouping) shares similar traits to what is implemented in the Oracle database [52]. However, the architectural differences with respect to a MapReduce system like Hadoop, and our quest to explore the design space and trade-offs of Rollup aggregates make such work complementary to ours.

4.1.3 Problem Statement

We now define the Rollup operation as a generalization of the SQL Rollup clause, introducing it by way of a running example. We use the same example in Section 4.1.4 to elucidate the details of design choices and, in Section 4.1.5, to benchmark our results.

Rollup can be thought of as a hierarchical Group By at various granularities, where the grouping keys at a coarser granularities are a subset of the keys at a finer granularity. More formally, we define the Rollup operation on an input *data set*, an *aggregation function*, and a set of n *hierarchical granularities*:

- We consider a *data set* akin to a database table, with M columns $c_1, \ldots, c_M$ and L rows $r_1, \cdots r_L$ such that each row r_i corresponds to the $(r_{i1}, \ldots, r_{iM})$ tuple.
- Given a set of rows $R \subseteq \{r_1, \cdots r_L\}$, an *aggregation function* $f(R)$ produces our desired result.
- n *granularities* $d_1, \ldots, d_n$ determine the groupings that an input data is subject to. Each d_i is a subset of $\{c_1, \cdots c_M\}$, and granularities are *hierarchical* in the sense that $d_i \subsetneq d_{i+1}$ for each $i \in [1, n-1]$.

The Rollup computation returns the result of applying f after grouping the input by the set of columns in each granularity. Hence, the output is a new table with tuples corresponding to grouping over the finest (d_n) up to the coarsest (d_1) granularity,

denoting irrelevant columns with an ALL value [64].

Example. Consider an Internet Service provider which needs to compute aggregate traffic load in its network, per day, month, year and in overall. We assume input data to be a large table with columns (c_1, c_2, c_3, c_4) corresponding to (year, month, day, payload[2]). A few example records from this dataset are shown in the following:

```
(2012,  3, 14, 1)
(2012, 12,  5, 2)
(2012, 12, 30, 3)
(2013,  5, 24, 4)
```

The aggregation function f outputs the sum of values over the c_4 (payload) column. Besides SUM, other typical aggregation functions are MIN, MAX, AVG and COUNT; it is also possible to consider aggregation functions that evaluate data in multiple columns, such as for example correlation between values in different columns.

Input granularities are $d_1 = \emptyset$, $d_2 = \{\text{year}\}$, $d_3 = \{\text{year}, \text{month}\}$, and $d_4 = \{\text{year}, \text{month}, \text{day}\}$. The highest granularity, $d_1 = \emptyset$, groups on no columns and is therefore equivalent to a SQL Group By ALL clause that computes the overall sum of the payload column; such an overall aggregation is always computed in SQL implementations, but it is not required in our more general formulation. We will see in the following that "global" aggregation is problematic in MapReduce.

In addition to aggregation on hierarchical time periods as in this case, Rollup aggregation applies naturally to other cases where data can be organized in tree-shaped taxonomies, such as for example country-state-region or unit-department-employee hierarchies.

If applied on the example, the Rollup operation yields the following result (we use '*' to denote ALL values):

```
(2012,  3, 14,  1)
(2012,  3,  *,  1)
(2012, 12,  5,  2)
(2012, 12, 30,  3)
(2012, 12,  *,  5)
(2012,  *,  *,  6)
(2013,  5, 24,  4)
(2013,  5,  *,  4)
(2013,  *,  *,  4)
(   *,  *,  *, 10)
```

[2] In Kilobytes

Rows with ALL values represent the result of aggregation at coarser granularities: for example, the `(2012, *, *, 6)` tuple is the output of aggregating all tuples from year 2012.

Aggregation Functions and Combiners. In MapReduce, it is possible to pre-aggregate values computed in mappers by defining *combiners*. We will see in the following that combiners are crucial for the performance of algorithms defined in MapReduce. While many useful aggregation functions are subsceptible to being optimized through combiners, not all of them are. Based on the definition by Gray *et al.* [64], when an aggregation function is *holistic* there is no constant bound on the size of a combiner output; representative holistic functions are MEDIAN, MODE and RANK.

The algorithms we define are differently subsceptible to the presence and effectiveness of combiners. When discussing the merits of each implementation, we also consider the case where aggregation functions are holistic and hence combiners are of little or no use.

4.1.4 The design space

We explore the design space of Rollup, with emphasis on the trade-off between communication cost and parallelism. We first apply a model to obtain theoretical bounds on replication rate and reducer key size; we then consider two algorithms (Vanilla and In-Reducer Grouping) that are at the end-points of the aforementioned trade-off, having respectively maximal parallelism and minimal communication cost. We follow up by proposing various algorithms that operate in different, and arguably more desirable, points of the trade-off space.

4.1.4.1 Bounds on Replication and Parallelism

Here we adopt the model by Afrati *et al.* [49] to find upper and lower bounds for the replication rate. Note that the model, unfortunately, does not account for combiners nor for multi-round MapReduce algorithms.

First, we define the number of all possible inputs and outputs to our problem, and a function $g(q)$ that allows to evaluate the number of outputs that can be covered with i input records. To do this, we refer to the definitions in Section 4.1.3:

1. **Input set**: we call C_i the number of different values that each column c_i can take. The total number of inputs is therefore $|I| = \prod_{i=1}^{M} C_i$.

2. **Output set**: for each granularity d_i, we denote the number of possible grouping keys as $N_i = \prod_{C_i \in d_i} C_i$ and the number of possible values that the aggregation function can output as A_i.[3] Thus, the total number of outputs is $|O| = \sum_{i=1}^{n} N_i A_i$.
3. **Covering function**: let us consider a reducer that receives q input records. For each granularity d_i, there are N_i grouping keys, each one grouping $|I|/N_i$ inputs and producing A_i outputs. The number of groups that the reducer can cover at granularity d_i is therefore no more than $\lfloor qN_i/|I| \rfloor$, and the covering function is $g(q) = \sum_{i=1}^{n} A_i \left\lfloor \frac{qN_i}{|I|} \right\rfloor$.

Lower Bound on Replication Rate. We consider p reducers, each receiving $q_i \leq q$ inputs and covering $g(q_i)$ outputs. Since together they must cover all outputs, it must be the case that $\sum_{j=1}^{p} g(q_j) \geq |O|$. This corresponds to

$$\sum_{j=1}^{p} \sum_{i=1}^{n} A_i \left\lfloor \frac{q_j N_i}{|I|} \right\rfloor \geq \sum_{i=1}^{n} N_i A_i. \tag{4.1}$$

Since $q_j N_i/|I| \geq \lfloor q_j N_i/|I| \rfloor$, we obtain the lower bound of the replication rate r as:

$$r = \sum_{i=1}^{p} \frac{q_i}{|I|} \geq 1. \tag{4.2}$$

Equation 4.2 seems to imply that Rollup aggregates is an *embarassingly parallel* problem: the $r \geq 1$ bound on replication rate does not depend on the size q_i of reducers. In Section 4.1.4, we show – for the first time – an instance of an algorithm that matches the lower bound. Instead, known instances of Rollup aggregates have a larger replication rate, as we shall see next.

Limits on Parallelism. Let us now reformulate Equation 4.2, this time requiring only that the output of the coarsest granularity d_1 is covered. We obtain

$$\sum_{j=1}^{p} \left\lfloor \frac{q_j N_1}{|I|} \right\rfloor \geq N_1.$$

Clearly, the output *cannot be covered* (the left side of the equation would be zero) unless there are reducers receiving at least $q_j \geq |I|/N_1$ input records. Indeed, the coarsest granularity imposes hard limits on the parallelism, requiring to broadcast the full input on at most N_1 reducers. This is exacerbated if – as it is the case with the standard SQL Rollup – there is an overall aggregation, resulting in $d_1 = \emptyset$, $N_1 = 1$ and therefore

[3] For the limit case $d_i = \emptyset$, $N_i = 1$, corresponding to the single empty grouping key.

$q_j \geq |I|$. A *single reducer* needs to receive *all the input*: it appears that no parallelism whatsoever is possible.

As we show in the following, this negative result however depends on the limitations of the model: by applying combiners and/or multiple rounds of MapReduce computation, it is indeed possible to compute efficient Rollup aggregates in parallel.

Maximum Achievable Parallelism. Our model considers parallelism as determined by the number of reducers p and the number of input records q_j each of them processes. However, one may also consider the number of *output* records produced by each reducer: in that case, the maximum parallelism achievable is when each reducer produces at most a single output value. This can be obtained by assigning each grouping key in each granularity to a different reducer; the aggregation function is then guaranteed to output only one of the A_i possible values. This, however, implies a replication rate $r = n$; an implementation of the idea is described in the following section.

4.1.4.2 Baseline algorithms

Next, we define two baseline algorithms to compute Rollup aggregates: Vanilla, which is discussed in [69], and *In Reducer Grouping*, which is our contribution. Then, we propose a hybrid approach that combines both baseline techniques.

Vanilla Approach.

We describe here an approach that maximizes parallelism at the detriment of communication cost; since this is the approach which is currently implemented in Apache Pig [67] we refer to it as Vanilla. Nandi *et al.* [69] refer to it as "naive".

The Rollup operator can be considered as the result multiple Group By operations: each of them is carried out at a different granularity. Thus, to perform Rollup on *n* granularities, for each record, the vanilla approach generates exactly *n* records corresponding to these *n* grouping sets (each grouping sets belongs to one granularity). For instance, taking as input the `(2012, 3, 14, 1)` record of the running example, this approach generates 4 records as outputs of the map phase:

```
(2012, 3, 14, 1) (day granularity)
(2012, 3,  *, 1) (month granularity)
(2012, *,  *, 1) (year granularity)
(  *, *,  *, 1) (overall granularity)
```

The Reduce step performs exactly as the reduce step of a Group By operation, using the first three records (year, month, day) as keys. By doing this, reducers pull all the data that is needed to generate each output record (shuffle step), and compute the aggregate

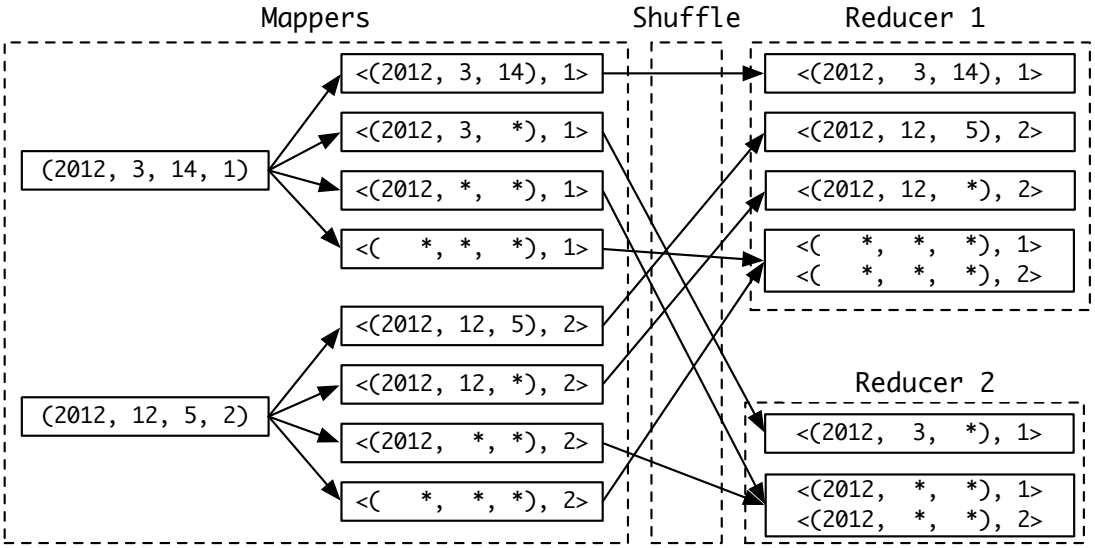

Figure 4.1: Example for the vanilla approach.

(reduce step). Figure 4.1 illustrates a walk-through example of the vanilla approach with just 2 records.

Parallelism and Communication Cost. The final result of Rollup is computed in a single MapReduce job. As discussed above, this implementation obtains the maximum possible degree of parallelism, since it can be parallelized up to a level where a single reducer is responsible of a single output value. On the other hand, this algorithm requires maximal communication costs, since for each input record, n map output records are generated. In addition, when the aggregation operation is *algebraic* [64], redundant computation is carried out in the reduce phase, since results computed for finer granularities cannot be reused for the coarser ones.

Impact of Combiners. This approach largely benefits from combiners whenever they are available, since they can compact the output computed at the coarser granularity (*e.g.*, in the example the combiner is likely to compute a single per-group value at the *year* and overall granularity). Without combiners, a granularity such as overall would result in shuffling data from *every input tuple* to a *single reducer*.

While combiners are very important to limit the amount of data sent along the network, the large amount of temporary data generated with this approach is still problematic: map output tuples need to be buffered in memory, sorted, and eventually spilled to disk if the amount of generated data does not fit in memory. This results, as we show in Section 4.1.5, in performance costs that are discernible even when combiners are present.

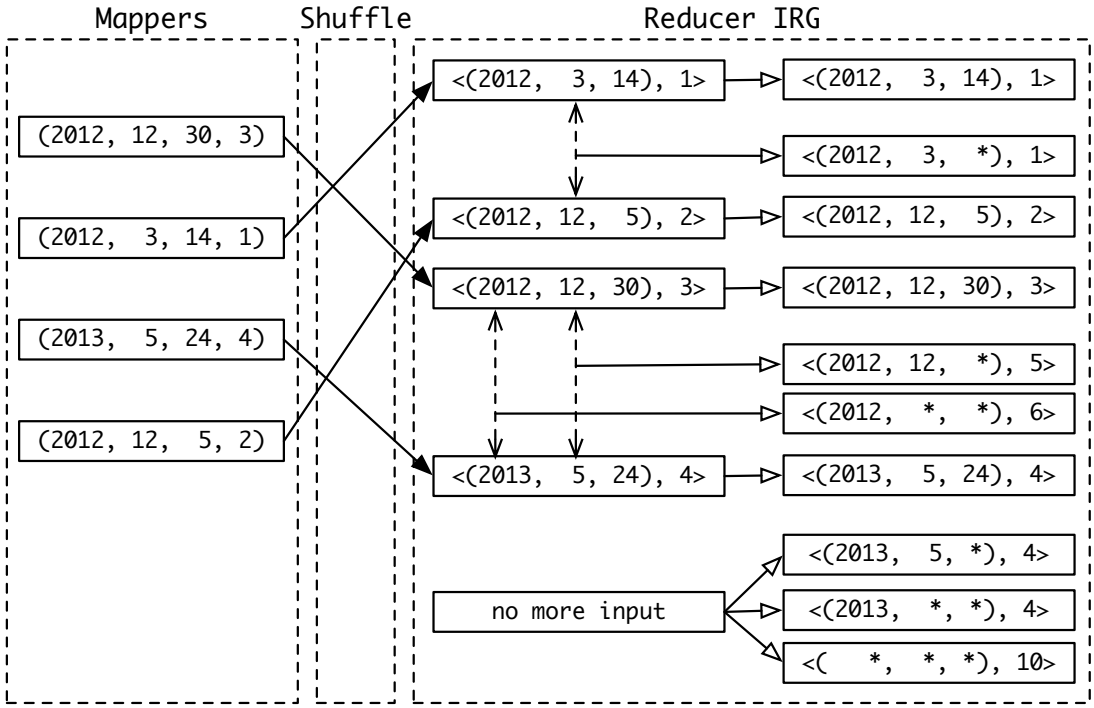

Figure 4.2: Example for the IRG approach.

In-Reducer Grouping. After analyzing an approach that maximizes parallelism, we now move to the other end of the spectrum and design an algorithm that minimizes communication costs. In contrast to the Vanilla approach, where the complexity resides on the Map phase and the Reduce phase behaves as if implementing an ordinary Group By clause, we propose an In-Reducer Grouping (IRG) approach, where all the logic of grouping is performed in the Reduce phase.

In-Reducer Grouping makes use of the possibility to define a *partitioner* in Hadoop [60, 75]. The mapper selects the columns of interest (in our example, all columns are needed, so the map function is simply the identity function). The keys are the finest granularity d_n (*day* in our example) but data is partitioned only by the columns of the coarsest granularity d_1. In this way, we can make sure that *1)* each reducer receives enough data to compute the aggregation function even for the coarsest granularity d_1; *2)* the intermediate keys are sorted [60, 75], so for every grouping key of any granularity d_i, the reducer will process consecutively *all* records pertaining to the given grouping key.

Figure 4.2 shows an example of the IRG approach. The mapper is the identity function, producing *(year, month, day)* as the keys and *payload* as the value. The coarsest granularity d_1 is overall, and $N_1 = 1$: hence, all $\langle \text{key}, \text{value} \rangle$ pairs are sent to a single reducer. The reducer groups all values of the same key, and processes the list of values associated to that key, thus computing the sum of all values as the total payload t. The grouping

logic in the reducer also takes care of sending t to n grouping keys constructed from the reducer input key. For example, with reference to Figure 4.2, the input pair `(<2012, 3, 14>, 1)` implies that value $t = 1$ is sent to grouping keys `(2012, 3, 14)`, `(2012, 3, *)`, `(2012, *, *)` and `(*, *, *)`. The aggregators in these grouping keys accumulate all t values they receive. When there is no more t value for a grouping key (in our example, when *year* or *month* change, as shown by the dashed lines in Figure 4.2), the aggregator outputs the final aggregated value.

The key observation we exploit in the IRG approach is that a secondary, lexicographic sorting, is fundamental to minimize state in the reducers. For instance, at *month* granularity, when the reducer starts processing pair `(<2013, 5, 24>, 4)`, then we are guaranteed that all grouping keys of month smaller than `(2013, 5)` (*e.g.* `(2012, 12)`) have already been processed and should be output without further delay. This way reducers need not keep track of aggregators for previous grouping keys: reducers only use n aggregators, one for each granularity.

To summarize, the IRG approach extensively relies on the idea of an *on-line algorithm*: it makes a single pass over its input, maintaining only the necessary state to accumulate aggregates (both algebraic and holistic) at different granularities, and produces outputs as the reduce function iterates over the input.

Parallelism and Communication Cost. Since mappers output one tuple per input record, the replication rate of the IRG algorithm meets the lower bound of 1, as showed in Equation 4.2. On the other hand, this approach has limited parallelism, since it uses no more reducers than the number N_1 of grouping keys at granularity d_1. In particular, when an overall aggregation is required, IRG can only work on a *single reducer*. As a result, IRG is likely to perform less work and require less resources than the Vanilla approach described previously, but it cannot benefit from parallelization in the reduce phase.

Impact of Combiners. Since the IRG algorithm minimizes communication cost, combiners only perform well if pre-aggregation at the finest granularity d_n is beneficial – *i.e.*, if the number of rows L in the data set is definitely larger than the number of grouping keys at the finest granularity, N_n. As such, the performance of the IRG approach suffers the least from the absence of combiners, *e.g.* when aggregation functions are not algebraic.

If the aggregate function is algebraic, however, the IRG algorithm is designed to re-use results from finer granularities in order to build the aggregation function *hierarchically*: in our running example, the aggregate of the total payload processed in a month can be obtained by summing the payload processed in the days of that month, and the aggregate for a year can likewise be computed by adding up the total payload for each month. Such an approach saves and reuses computation in a way that is not possible to obtain with the Vanilla approach.

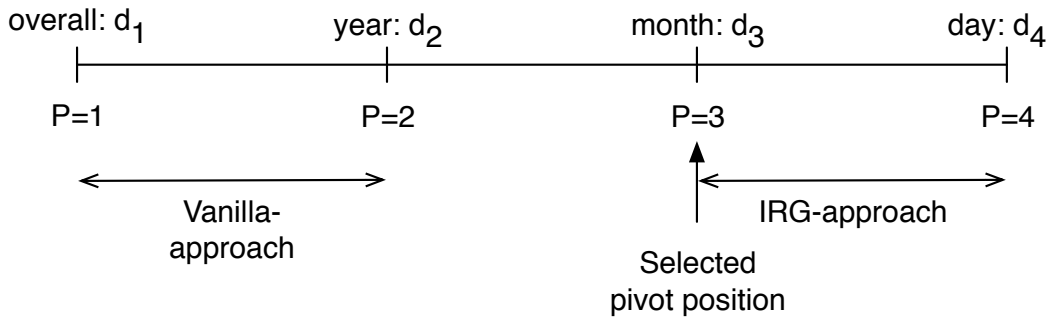

Figure 4.3: Pivot position example.

Hybrid approach: Vanilla + IRG. We have shown that Vanilla and IRG are two "extreme" approaches: the first one maximizes parallelism at the expense of communication cost, the second one instead minimizes communication cost but does not provide good parallelism guarantees.

Neither approach is likely to be an optimal choice for such a tradeoff: in a realistic system, we are likely to have way less reducers than number of output tuples to generate (therefore making the extreme parallelism guarantees produced by Vanilla excessive); however, in particular when an overall aggregate is needed, it is reasonable to require an implementation that does not have the bottleneck of a single reducer.

In order to benefit at once from an acceptable level of parallelism and lower communication overheads, we propose an *hybrid* algorithm that fixes a *pivot* granularity P: all aggregate functions on granularities between d_P and d_n are computed using the IRG algorithm, while aggregates for granularities above d_P are obtained using the Vanilla approach. A choice of $P = 1$ is equivalent to the IRG algorithm, while $P = n$ corresponds to the Vanilla approach.

Let us consider again our running example, and fix the pivot position at $P = 3$, as shown in Figure 4.3. This choice implies that aggregates for the overall and *year* granularities d_1, d_2 are computed using the Vanilla approach, while aggregates for the other granularities d_3, d_4 (*month* and *day*) are obtained using the IRG algorithm. For example, for the `(2012, 3, 14, 1)` tuple, the hybrid approach produces *three* output records at the mapper:

```
(2012, 3, 14, 1) (day granularity)
(2012, *,  *, 1) (year granularity)
(   *, *,  *, 1) (overall granularity)
```

In this case the map output key space is partitioned by the month granularity, so that there is *1)* one reducer per each month in the input dataset, that computes aggregates for granularities up to the month level, and *2)* multiple reducers that compute aggre-

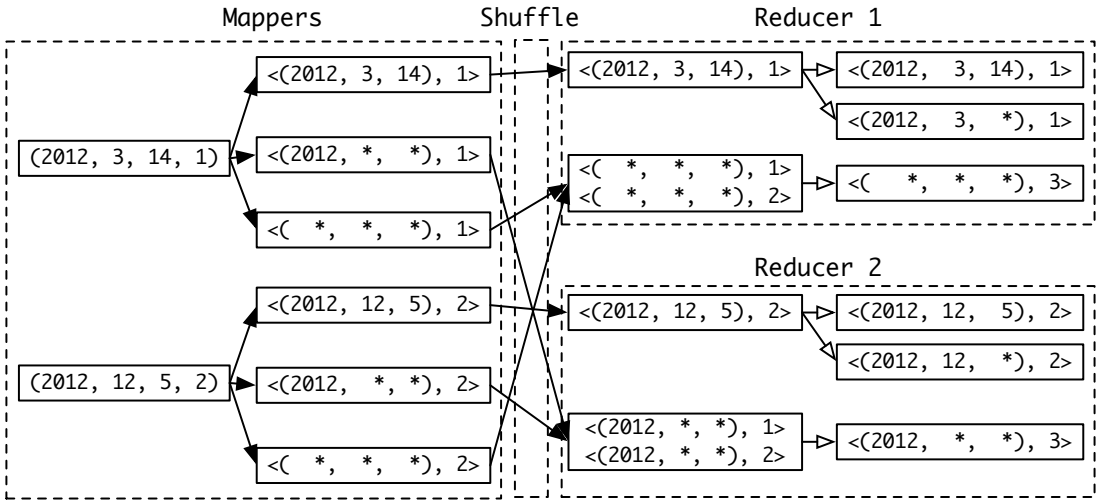

Figure 4.4: Example for the Hybrid Vanilla + IRG approach.

gates for the *overall* and *year* granularities. Figure 4.4 illustrates an example with two reducers.

Some remarks are in order. Assuming a uniform distribution of the input dataset, the load on reducers of type *1)* is expected to be evenly shared, as an input partition corresponds to an individual month and not the whole dataset. The load on reducers of type *2)* might seem still prohibitive; however, we note that when combiners are in place they are going to vastly reduce the amount of data sent to the reducers responsible of the overall and *year* aggregate computation. For our example, the reducers of type *2)* receive few input records, because the overall and *year* aggregates can be largely computed in the map phase. Furthermore, we remark that the efficiency of combiners in reducing input data to reducers (and communication costs) is very high for coarse granularities, and decreases towards finer granularities: this is why the IRG algorithm applies the Vanilla approach from the pivot position, up to coarse granularities.

Parallelism and Communication Cost. The performance of the hybrid algorithm depends on the choice of P: the replication rate (before combiners) is P. The number of reducer that this approach can use is the total of *1)* N_P grouping keys that are handled with the IRG algorithm, and *2)* $\sum_{i=0}^{P-1} N_i$ grouping keys that are handled with the Vanilla approach. Ideally, an *a priori* knowledge of the input data can be used to guide the choice of the pivot position. For example, if the data in our running example is known to span over tens of years and we know we only have ten reducer slots available (*i.e.*, at most ten reducer tasks can run concurrently), a choice of partitioning by year ($P = 2$) would be reasonable. Conversely, if the dataset only spans a few years and hundreds of reducer slots are available, then it would be better to be more conservative and choose

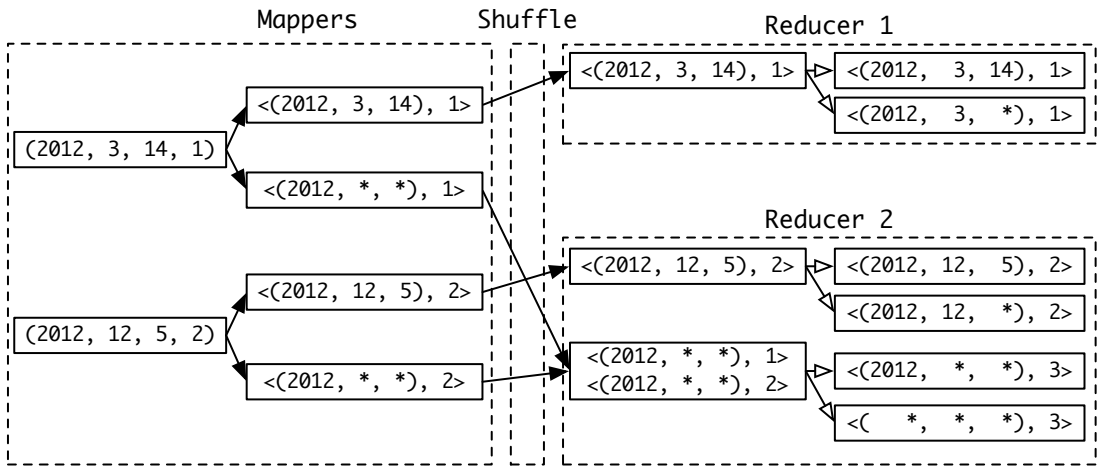

Figure 4.5: Example for the Hybrid IRG + IRG approach.

$P = 3$ or $P = 4$ to obtain better parallelism at the expense of a higher communication cost.

Impact of Combiners. The hybrid approach heavily relies of combiners. Indeed, when combiners are not available, all input data will be sent to the one reducer in charge of the overall granularity; in this case, it is then generally better to choose $P = 1$ and revert to the IRG algorithm. However, when the combiners are available, the benefit for the hybrid approach is considerable, as discussed above.

4.1.4.3 Alternative hybrid algorithms

We now extend the hybrid approach we introduced previously, and propose two alternatives: a single job involving two parallel IRG instances, and a chained job involving a first IRG computation and a final IRG aggregation.

Hybrid approach: IRG + IRG. In the previous section, we have shown that it is possible to design an algorithm aiming at striking a good balance between parallelism and replication rate, using a single parameter, *i.e.* the pivot position. In the baseline hybrid approach, parallelism is an increasing function of the replication rate, so that better parallelism is counterbalanced by higher communication costs in the shuffle phase.

Here, we propose an alternative approach that results in a constant replication rate of 2: the "trick" is to replace the Vanilla part of the baseline hybrid algorithm with a second IRG approach. Using the same running example as before, for the tuple `(2012, 3, 14, 1)`, and selecting the pivot position $P = 3$, the *two* map output records are:

```
(2012, 3, 14, 1) (day granularity)
(2012, *, *, 1) (year granularity)
```

Figure 4.5 on the facing page illustrates a running example. In this case, the map output key space is partitioned by the *month* granularity, such that there is one reducer per month that uses the IRG algorithm to compute aggregates; in addition, there is *one reducer* receiving all tuples having ALL values taking care of the *year* and overall granularities, using again the IRG approach. As before, the role of combiners is crucial: the amount of (year, *, *, payload) tuples that are sent to the single reducer taking care of *year* and *overall* aggregates is likely to be very small, because opportunities to compute partial aggregates in the map phase are higher for coarser granularities.

Parallelism and Communication Cost. This algorithm has a constant replication rate of 2. As we show in Section 4.1.5, the choice of the pivot position P is here much less decisive than for the baseline hybrid approach: this can be explained by the fact that moving the pivot to finer granularities does not increase communication costs, as long as the load on the reducer taking care of the aggregates for coarse granularities remains low.

Impact of Combiners. Similarly to the baseline hybrid approach, this algorithm relies heavily on combiners; if combiners are not available, then, a simple IRG approach would be preferable.

Chained IRG. It is possible to further decrease the replication rate and hence the communication costs of computing Rollup aggregates by adopting a multi-round approach composed of two chained MapReduce jobs. In this case, the first job pre-aggregates results up to the pivot position P using the IRG algorithm; the second job uses partial aggregates from the first job to produce – on a single reducer – the final aggregate result, again using IRG. We note here that a similar observation, albeit for computing matrix multiplication, is also discussed in detail in [49].

Parallelism and Communication Cost. The parallelism of the first MapReduce job is determined by the amount N_P of grouping keys at the pivot position; the second MapReduce job, has a single reducer. However, the input size of the second job is likely to be orders of magnitude smaller than the first one, so that the runtime of the reduce phase of the second job – unless the pivot position puts too much effort on the second job – is generally negligible. The fact that the second reducer operates on a very small amount of input, results in a replication rate very close to 1.

The main drawback of the chained approach is due to job scheduling strategies: if jobs are scheduled in a system with idle resources, as we show in Section 4.1.5, the chained IRG algorithm results in the smallest runtime. However, in a loaded system, the second (and in general very small) MapReduce job could be scheduled later, resulting in artificiously large delays between job submission and its execution.

Impact of Combiners. This approach does not rely heavily on combiners *per se*. However, it requires the aggregation function to be algebraic in order to make it possible for the second MapReduce job to re-use partial results.

4.1.5 Experimental Evaluation

We now proceed with an experimental approach to study the performance of the algorithms we discussed in this work. We use two main metrics: *runtime* – *i.e.* job execution time – and *total amount of work*, *i.e.* the sum of individual task execution times. Runtime is relevant on idle systems, in which job scheduling does not interfere with execution times; total amount of work is instead an important metric to study in heavily loaded systems where spare resources could be assigned to other pending jobs.

4.1.5.1 Experimental Setup

Our experimental evaluation is done on a Hadoop cluster of 20 slave machines (8GB RAM and a 4-core CPU) with 2 map and 1 reduce slot each. The HDFS block size is set to 128MB. All results shown in the following are the average of 5 runs: the standard deviation is smaller than 2.5%, hence – for the sake of readability – we omit error bars from our figures.

We compare the five approaches described in Section 4.1.4: baseline algorithms (Vanilla, IRG, Hybrid Vanilla + IRG) and alternative hybrid approaches (Hybrid IRG + IRG Chained IRG). We evaluate a single Rollup aggregation job over (*overall*, *year*, *month*, *day*, *hour*, *minute*, *second*) that uses the SUM aggregate function which, being algrebraic, can benefit from combiners. Our input dataset is a synthetic log-trace representing historical traffic measurements taken by an Internet Service Provider (ISP): each record in our log has *1)* a time-stamp expressed in (*year*, *month*, *day*, *hour*, *minute*, *second*); and *2)* a number representing the payload (e.g. number of bytes sent or received over the ISP network). The time-stamp is generated uniformly at random within a variable number of years (where not otherwise specified, the default is 40 years). The payload is a uniformly random positive integer. Overall, our dataset comprises 1.5 billion binary tuples of size 32 bytes each, packed in a SequenceFile [75].

4.1.5.2 Results

This section presents a range of results we obtained in our experiments. Before delving into a comparative study of all the approaches outlined above, we first focus on studying the impact of combiners on the performance of the Vanilla approach. Then, we move to a detailed analysis of runtime and amount of work for baseline algorithms (Vanilla,

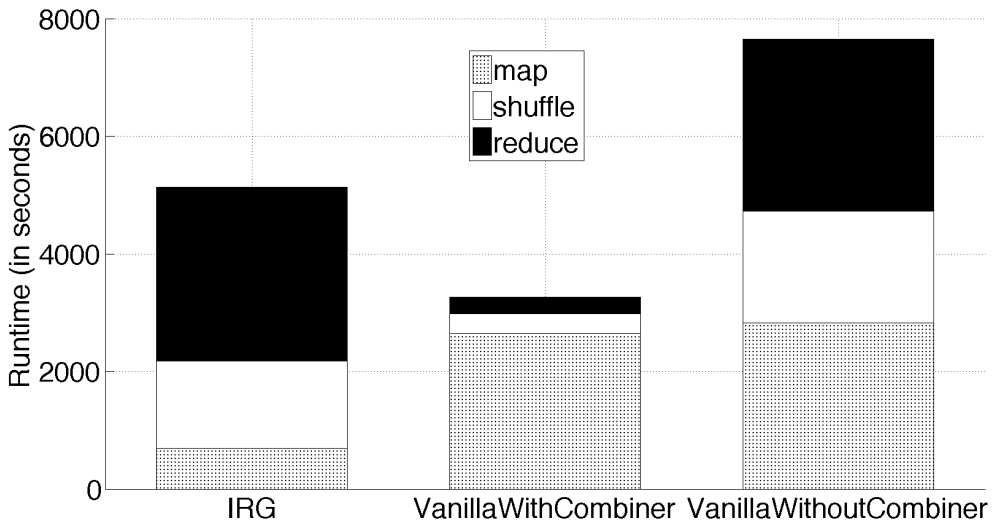

Figure 4.6: Impact of combiners on runtime for the Vanilla approach.

IRG, and Hybrid), and we conclude with an overview to outline merits and drawbacks of alternative hybrid approaches.

The role of combiners. Figure 4.6 illustrates a break-down of the runtime for computing the Rollup aggregate on our dataset, showing the time a job spend in the various phases of a MapReduce computation. Clearly, combiners play an important role for the Vanilla approach: they are beneficial in the shuffle and reduce phases. When combiners cannot be used (e.g. because the aggregation function is not algebraic), the IRG algorithm outperforms the Vanilla approach. With combiners enabled, the IRG algorithm is slower (larger runtimes) than the Vanilla approach: this can be explained by the lack of parallelism that characterizes IRG, wherein a single reducer is used as opposed to 20 reducers for the Vanilla algorithm. Note that, in the following experiments, combiners are systematically enabled. Finally, Figure 4.6 confirms that the IRG approach moves algorithmic complexity from the map phase to the reduce phase.

Baseline algorithms. In Figure 4.7 we compare the runtime of Vanilla, IRG, and the hybrid Vanilla + IRG approach. In our experiments we study the impact of the pivot position P, in lights of the "nature" of the input dataset: we synthetically generate data such that they span 1, 10 and 40 years worth of traffic logs.[4]

Clearly, IRG (which corresponds to $P = 1$) is the slowest approach in terms of runtime. Indeed, using a single reducer incurs in prohibitive I/O overheads: the amount of

[4]Note that the size – in terms of number of tuples – of the input data is kept constant, irrespectively of the number of represented years.

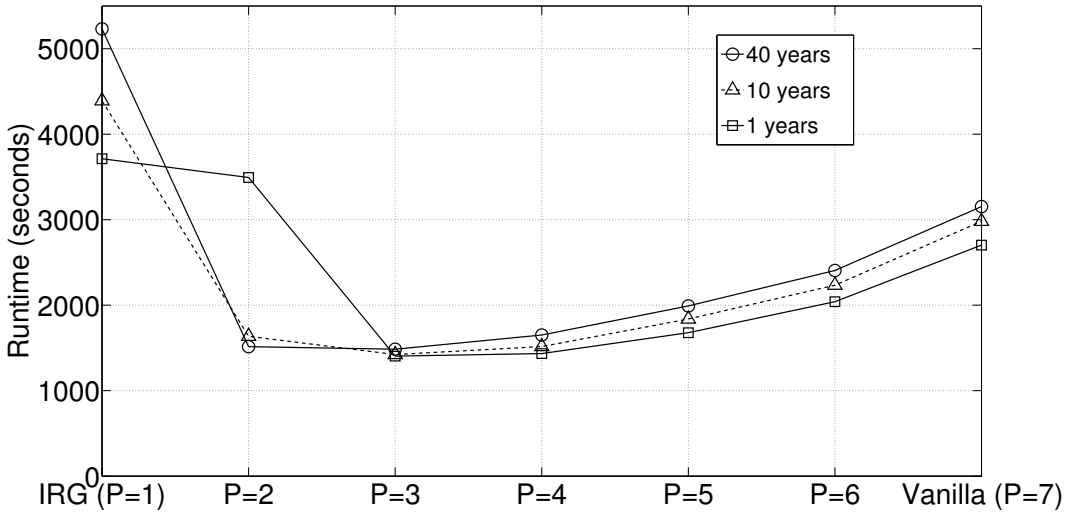

Figure 4.7: Runtime comparison of baseline approaches.

data shuffled into a single reducer is too large to fit into memory, therefore spilling and merging operations at the reducer proceed at disk speeds. Although no redundant computations are carried out in IRG, I/O costs outweigh the savings in computations.

A hybrid approach ($2 \leq P \leq 6$) outperforms both IRG and Vanilla algorithms, with runtime as little as half that of the Vanilla approach. Communication costs make the runtime grow slowly as the pivot position moves towards finer granularities, suggesting that in doubt, it is better to position the pivot to the right (increased communication costs) rather than to the left (lack of parallelism). In our case, where a maximum of 20 reduce tasks can be scheduled at any time, our results indicate that P should be chosen such that N_P is larger than the number of available reducers. As expected, experiments with data from a single year indicate that the pivot position should be placed further to the right: the hybrid approach with $P = 2$ essentially performs as badly as the single-reducer IRG.

Now, we present our results under a different perspective: we focus on the *total amount of work* executed by a Rollup aggregate implemented according to our baseline algorithms. We define the total amount of work for a given job as the sum of the runtime of each of its (map and reduce) tasks. Figure 4.8 indicates that the IRG approach consumes the least amount of work. By design, IRG is built to avoid redundant work: it has minimal replication rate, and the single reducer can produce Rollup aggregates with a single pass over its input.

As a general remark, that applies to all baseline algorithms, we note that the total amount of work is largely determined by the map phase of our jobs. The trend is tangible as P moves toward finer granularities: despite communication costs (the shuffle phase, which

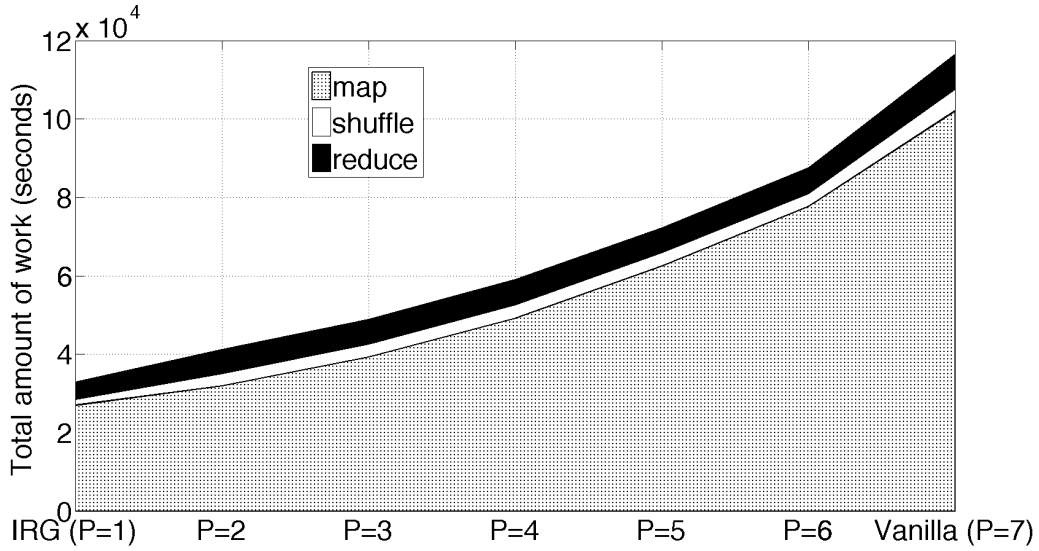

Figure 4.8: Amount of work comparison of baseline approaches.

accounts for the replication rate) do not increase much with higher values of P thanks to the key role of combiners, map tasks still need to materialize data on disk before it can be combined and shuffled, thus contributing to a large extent to higher amounts of work.

Alternative Hybrid Approaches. We now give a compact representation of our experimental results for variants of the Hybrid approach we introduce in this work. Figure 4.9 offers a comparison, in terms of job runtime, of the Hybrid Vanilla + IRG approach to the Hybrid IRG + IRG and the Chained IRG algorithms. For the sake of readability, we omit from the figure experiments corresponding to $P = 1$ and $P = 7$.

Figure 4.9 shows that the job runtime of the Hybrid Vanilla + IRG algorithm is sensitive to the choice of the pivot position P. Despite the use of combiners, the Vanilla "component" of the hybrid algorithm largely determines the job runtime, as discussed above. The IRG + IRG hybrid algorithm obtains lower job runtime and is less sensitive to the pivot position, albeit $3 \leq P \leq 5$ constitutes an ideal range in which to place the pivot. The best performance in terms of runtime is achieved by the Chained IRG approach: in this case, the amount of data shuffled through the network (aggregated over each individual job of the chain) is smaller than what can be achieved by a single MapReduce job. We further observe that placing P towards finer granularities contributes to small job runtime: once an appropriate level of parallelism can be achieved in the first job of the chain, the computation cost of the second job in the chain is negligible, and the total amount of work (not shown here due to space limitations) is almost constant and extremely close to the one for IRG.

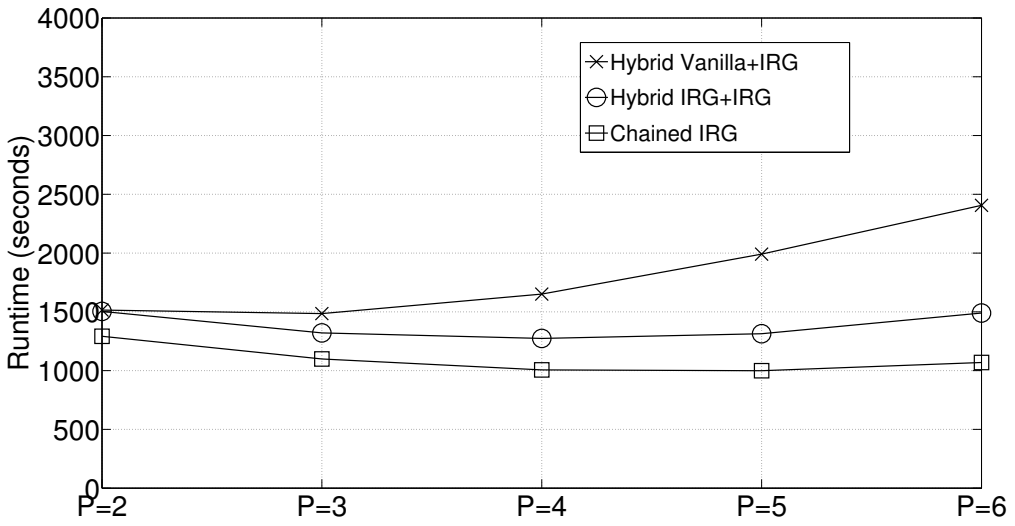

Figure 4.9: Comparison between alternative hybrid approaches.

We can now summarize our findings as follows:

- All the approaches that we examined greatly benefit from the, fortunately common, property that aggregation functions are algebraic and therefore enable combiners and re-using partial results. If this is not the case, approaches based on the IRG algorithm are preferable.
- If total amount of work is the metric to optimize, IRG is the best solution because it minimizes redundant work. If low latency is also required, hybrid approaches offer a good trade-off, provided that the pivot position P is chosen appropriately.
- Our alternative hybrid approaches are the best performing solutions; both are very resilient to bad choices of the P pivot position, which can therefore be chosen with a very rough a-priori knowledge of the input dataset. Chained IRG provides the best results due to its minimal communication costs. However, chained jobs may suffer from bad scheduling decisions in a heavily loaded cluster, as the second job in the chain may "starve" due to large jobs being scheduled first. The literature on MapReduce scheduling offers solutions to this problem [27].

4.2 Efficient and Self-Balanced Rollup Operator for MapReduce-like Systems

4.2.1 Introduction

Despite the importance of data aggregation, the field of large-scale data-intensive computing systems – where data can reach petabytes and be distributed on clusters of thousands of nodes – has not seen much effort toward efficient implementations of the above concepts. In the Hadoop MapReduce ecosystem [1,60], high-level query languages such as Pig Latin [8] and HiveQL [7] offer simple implementations of the above constructs, which do not perform aggressive optimizations. In enterprise workloads, jobs coming from queries written in high-level languages are the majority [71]; an optimized implementation of these operators is therefore truly desirable.

Our work in the previous Section 4.1 proposes several efficient physical algorithms for MapReduce-like Rollup. In this Section, we focus on the design and implementation of a light-weight, cost-based optimization module for Rollup for high-level languages. We call this module: a *Rollup operator*.

Existing implementations are not satisfying: as we discuss in Section 4.2.2, current Rollup algorithms are naively biased toward extreme levels of parallelism. As a consequence, these approaches trade a theoretical possibility of scaling several orders of magnitude beyond the scale attainable by real-world clusters with a very significant overhead in terms of communications.

There are alternative Rollup algorithms (Section 4.1) that allow tuning the level of parallelism and the communication overhead of their implementation: with a proper setting, these algorithms perform better than naive implementations. Such approaches are appealing in the abstract, but they are practically very difficult for users to apply, since to determine the proper setting, they would require users to know: *(i)* the internals of the algorithm's implementation; *(ii)* statistical information about data distribution; *(iii)* size and performances of the cluster on which the algorithm is implemented. These requirements are difficult to meet at once, and they are essentially prohibitive in the context of high-level languages, whose very reason of existence is to hide this kind of complexities and allowing users to concentrate on extracting meaning from data.

As a result, in this Section, we design (in Section 4.2.3) and implement a new Rollup operator that is completely transparent to users. Our operator automatically collects statistics about data and cluster performance. Subsequently it uses this information to: *(i)* balance load across different nodes in the cluster; *(ii)* determine an appropriate operation point of Rollup algorithms using a lightweight cost-based optimizer.

We perform an extensive experimental campaign (Section 4.2.5) to assess the validity of our design choices, using both real and synthetic datasets, and comparing the performance of a variety of different Rollup algorithms. Results indicate that our Rollup operator, that relies on automatically tuned algorithms, delivers superior performance when compared to current Rollup implementations for the Apache Pig system. Our operator is released as open source software[5], and is currently in the process of being integrated in the upstream Apache Pig Latin language implementation. We summarize this work in Section 4.4.

4.2.2 Preliminaries & Related Work

Efficient computation of data summaries is an important topic that has received wide attention by the database community. Recently, such interest has led to several works to bring the benefits of data aggregation to MapReduce systems as well.

In this Section, we review several Rollup algorithms, for both traditional databases and MapReduce systems, and motivate the need for a substantially different approach. Rollup is a data operator first introduced by Gray *et al.* [64] as a special case of Cube. The Rollup operator aggregates data along a dimension by "climbing up" the dimension hierarchy. For example, to aggregate the volume of cars sold in the last several years, it is possible to use the Rollup operator to obtain sales along the time dimension, including multiple levels: `total`, `year`, `month`, and `day`. These *levels* form a hierarchy, of which `total` is the top level that includes aggregates from all records.

4.2.2.1 Rollup in Parallel Databases

Traditionally, the Rollup operator was studied through the lenses of its generalized operator, Cube, that groups data along all combinations (called *views*) of different levels in hierarchies. In Rollup, a view is equivalent to a level of the rollup hierarchy. In our car sales example, we have 4 different views.

Harinaryan *et al.* [66] introduced a model to evaluate the cost of executing Cube and a greedy algorithm to select a near-optimal execution plan. Agarwal *et al.* [51] proposed *top-down* algorithms (*PipeSort*, *PipeHash*) to optimize Cube computation: using finer group-bys to compute coarser ones (*i.e.* using `month` to compute `year`). Beyer and Ramakrishnan [58] suggested a *bottom-up computation* (BUC) to construct results from coarser to finer group-bys by reusing as much as possible the previously computed sort orders. All of these works are sequential algorithms which focus on single servers.

[5] https://bitbucket.org/bigfootproject/rollupmr

Scalable algorithms to handle the Rollup and Cube operators in parallel databases also received considerable attention. They are divided into two main groups: *work partitioning* [37, 48] and *data partitioning* [45, 46]. In work partitioning, each processor (or node) of the cluster computes aggregates for a set of one or many views independently. To do that, all processors access concurrently the entire dataset. Typically such an access is offered by a shared-disk array that is both expensive and difficult to scale in term of performance and size. Instead, data partitioning algorithms divide the input data set into various subsets. A node computes all views associated to the subsets of data it hosts. To obtain global aggregates, a subsequent *merge* phase is required. The main advantage of such a *Two Phase* algorithm is that nodes need not to have access to the whole data set but work on a small portion that can be easily stored on local memory/disks.

4.2.2.2 Rollup in MapReduce

In what follows, we assume the reader to be familiar with the MapReduce paradigm and its well-known open-source implementation Hadoop [60, 75].

Currently, MapReduce high-level languages such as Apache Pig and Hive borrow from parallel databases the Two Phase algorithm described previously. Its MapReduce variant is straightforward: each mapper computes aggregates of the whole Rollup hierarchy from its local data, then sends the partial results to reducers which merge and return the final aggregates. The MapReduce Two Phase (MRTP) algorithm, which can use combiners as an optimization, produces a large quantity of intermediate data that impose strain on network resources. In addition, a large number of intermediate data increases noticeably mapper overheads to sort and eventually spill them to disk. Similarly, since reducers compute aggregates in each view separately, the MRTP algorithm presents significant overheads due to redundant computation.

These overheads are the main reason for the MRTP inefficiency, and they can be avoided by employing the *in-reducer grouping* (IRG) design pattern [80]. IRG computes aggregates in a top-down manner by exploiting custom partitioning and sorting in MapReduce to move the grouping logic from the shuffle phase to reducers. However, IRG severely lacks parallelism: all processing is performed by a *single* reducer.

More recent works explore the design space of one-round MapReduce Rollup algorithms by studying the trade-off between communication cost (amount of data sent over network) and parallelism [28]. Through a parameter called pivot position, the Hybrid IRG+IRG algorithm (HII) provides users with the flexibility of choosing a sweet spot to balance parallelism and communication cost. For one-round MapReduce algorithms, the HII algorithm is shown (analytically and experimentally) to achieve the best performance, if and only if the pivot position is set to the optimal value. Otherwise, the HII

algorithm represent a valid theoretical contribution, albeit impractical to be used as a baseline for a high-level language operator.

For multi-round MapReduce Rollup algorithms, the work in [28] also proposes the ChainedIRG algorithm, which splits a Rollup operator into two chained MapReduce jobs based on a parameter (again, a pivot position). The pivot position divides the Rollup hierarchy between two jobs. The first job computes a subset of the hierarchy and the second job takes the first job's results to compute the rest.

Finally, MRCube, proposed by Nandi *et al.* [69], implements Cube and Rollup operators in three rounds. The first round performs record random sampling on input data and estimates the cardinality of reducer keys. It serves identifying large groups of keys whose cardinality exceeds the capacity of a reducer, and set a partition factor N. In the second round, MRCube splits these large groups into N sub-groups using *value partitioning*: two keys belong to the same sub-groups if and only if their values of the aggregated attribute are congruent modulo N. Each sub-group is computed as a partial aggregate by some reducer using the BUC algorithm [66]. The third round merges partial aggregates to produce final results. As a consequence, reducers do not handle excessive amounts of data, but the execution plan requires multi-rounds of MapReduce jobs. The authors also note that the value partitioning may be problematic when data is skewed on the aggregated attribute, as it can create sub-groups that exceed the reducer capacity and slow down the job. Also, as the number of large-groups may be high, the first step of MRCube can create a significant overhead and make MRCube slower than the MapReduce Two Phase algorithm [69].

4.2.2.3 Other Related Works

It is well known that MapReduce algorithms may suffer from poor performance if data is skewed. SkewReduce [53] is a framework that manages data and computation skew by using user domain knowledge to break the map and reduce tasks into smaller tasks; then it finds the optimal partition plan to achieve load balancing. SkewTune [54] is a dynamic skew mitigation system: by modifying the MapReduce architecture, it detects stragglers in reducers and pro-actively repartitions the unprocessed data to other idle reducers. Another approach is to design skew-resistant operators: data skew is handled at the algorithmic level. This approach does not require additional components, user interventions or modifications to the Hadoop framework. For instance, [56] proposes an algorithmic approach to handle set-similarity join. Apache Pig also supports a skew-resistant equi-join operator. For the MapReduce data aggregation, to the best of our knowledge, our work is the first to tackle a skew-resistant Rollup operator.

Finally, we consider related works that use a cost model to find optimal execution plans. SkewReduce [53] uses a cost model that requires two user-supplied cost functions. MR-

Share [57] proposes work-sharing optimizations based on a cost model to predict merged job's runtime. This model requires users to provide a set of constant, static parameters, that represent the underlying cluster performance. Both practices are not transparent to users. The latter is problematic in practice because cluster performance changes overtime, and in many cases, clusters are dynamically allocated (*e.g.* Amazon EC2, Google App Engine) which means unpredictable performance.

Instead, as we show in the next Section, our approach automatically measures cluster performance. These measurements are fed to a regression model that predicts the Rollup runtime. This is completely transparent to users, and adaptable to any cluster configuration.

4.2.3 A New Rollup Operator

In this Section, we describe our design of an efficient and skew-resistant Rollup operator. Our approach can be integrated directly in current MapReduce high-level languages such as Apache Pig and Hive, and it is completely transparent to users. Our design avoids any modifications to the Hadoop framework or the MapReduce programming model, thus making our work directly applicable to any MapReduce-like systems. Our approach can also handle a stateless design (*i.e.* no historical execution statistics), which is the current standard practice for systems that provide high-level languages on top of MapReduce.

4.2.3.1 Rollup Operator Design

Our Rollup operator has two main components: the *tuning job* and the *Rollup query*. The Rollup query can implement one of several Rollup algorithms, such as MRTP, HII, ChainedIRG, MRCube as described in Section 4.2.2.2. We discuss the choice of an appropriate algorithm in Section 4.2.3.2.

The heart of our work is the *tuning job*, a primary component with two main goals. The first is to determine how to achieve load-balancing taking into account skewed data when executing the Rollup query, which is clearly beneficial to any Rollup algorithm. The second goal is to determine the most suitable operating point of the Rollup query, provided that the underlying Rollup algorithm requires tuning. For example, Rollup algorithms like HII and ChainedIRG both rely on an essential parameter that, if not properly tuned, can lead to inefficient executions and bad performance. In this case, the tuning job automatically sets the parameter of such algorithms to the proper value, such that performance is maximized.

In addition, to obtain an efficient Rollup operator, our goal is to minimize the overhead caused by the tuning job. Consequently, we propose a single, light-weight tuning job that simultaneously carries out all the following tasks to produce efficient Rollup queries:

1. It produces representative samples of the input data;
2. It balances reducer loads using information on key distribution estimated from sample data;
3. If required, it determines appropriate parameters of a Rollup algorithms using a cost-based optimizer.

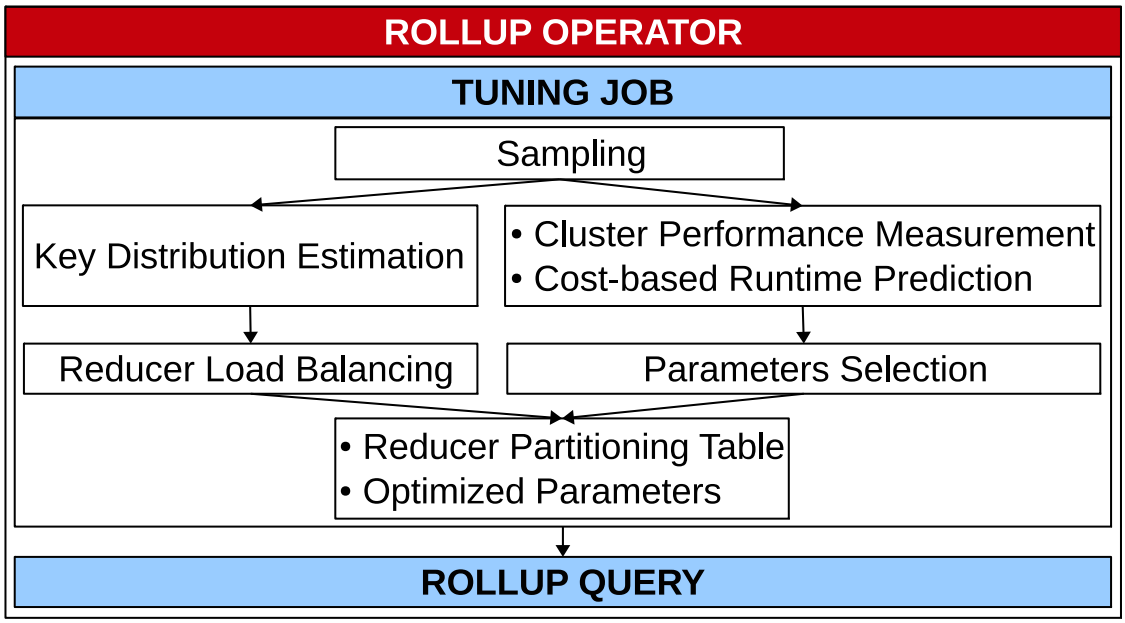

Figure 4.10: Overview of the Rollup operator design.

Figure 4.10 shows a sketch of our Rollup operator. The tuning job runs before the execution of the Rollup query. It produces a balanced partitioning table, and tunes parameters appropriately, when needed. The Rollup query uses these outcomes to optimize its performance.

In the rest of this Section, we describe in detail how our tuning job can fulfill its goals using a top-down presentation. First, we discuss our choice for the Rollup algorithm that implements the Rollup query in Section 4.2.3.2. Then, we present the internals of our tuning job in Section 4.2.3.3. We examine load balancing on reducers in Section 4.2.3.3.1, which uses statistics collected through sampling (Section 4.2.3.3.2). Finally, we present our cost model (Section 4.2.3.3.3) that steers the Rollup query toward an optimized operating point.

4.2.3.2 The Rollup Query: Algorithmic Choice

The design of our Rollup operator is generic enough to employ any Rollup algorithm. However, in this work, we focus on the two algorithms that proved to offer, consistently, superior performance when compared to alternative approaches, namely the Hybrid IRG+IRG (HII) and ChainedIRG [28]. Indeed, unlike MRTP, MRCube and IRG, such algorithms have the flexibility to adjust the level of parallelism to exploit, which translates into a much lower communication cost compared to MRTP and MRCube. Previous results in the literature [28] corroborate our choice, which we confirm in our extensive performance evaluation in Section 4.2.5.

Now, both ChainedIRG and HII require a fundamental parameter P called *pivot position*. Let us consider the Rollup dimension with n levels: $\{d_1, d_2, \ldots, d_n\}$ in which d_1 is the top-level of the hierarchy. P divides the Rollup dimensions in two subsets: $S_1 = \{d_1, \ldots, d_{P-1}\}$ and $S_2 = \{d_P, \ldots, d_n\}$. Each subset now represents a sub-Rollup query. The difference between ChainedIRG and HII is that, while HII computes each sub-Rollup query independently and has only one job, ChainedIRG exploits the aggregates on S_2 to compute aggregates on S_1 (hence, it requires two jobs). Nonetheless, both algorithms use the IRG design pattern to compute each sub-Rollup query.

Although the experimental results in [28] indicate that ChainedIRG has the best runtime in an isolated system, in this section we cast our Rollup operator on the HII algorithm because it is less prone to delays due to scheduling when the cluster is loaded. Instead, the runtime of multi-phase algorithms (such as ChainedIRG) could be inflated since the job scheduler can dedicate resources to other jobs in between the phases. Nonetheless, we note that our tuning job can easily accommodate alternative algorithms, and we show this in our experimental evaluation, where we present results of our Rollup operator with instances of the Rollup query implementing all the algorithms we discussed in Section 4.2.2.

Continuing our example in Section 4.2.2, if $P = 3$ the two subsets are $\{\texttt{total}, \texttt{year}\}$ and $\{\texttt{month}, \texttt{day}\}$. For each input record, the map phase of HII generates 2 $\langle \text{key}, \text{value} \rangle$ pairs of the bottom level of each subset (*i.e.* `year` and `day`).

Then for each subset, the mappers partition their $\langle \text{key}, \text{value} \rangle$ pairs by the top level of this subset (*i.e.* `total` and `month`). Taking advantage of this partitioning scheme and the sorting done by the MapReduce framework, day-to-month and year-to-total aggregates can be processed independently in the reduce phase. In fact, HII applies a top-down approach: we compute aggregates for each day in a month; then combine results to obtain month aggregates. Similarly, we compute results for each year and construct the total aggregate.

Choosing an appropriate pivot position P is crucial in HII and ChainedIRG, because it determines the parallelism as well as the communication cost of such algorithms. Values

of P that are too high can impose an excessive load on the reducer that is responsible for the `total` aggregates, and lessen the benefit of combiners (*i.e.* higher communication cost). Instead, low values of P could result in insufficient parallelism causing poor load balancing. Finding the proper value for P is a non-trivial problem.

4.2.3.3 The Tuning Job

We now discuss our main contribution, the tuning job, that balances the reducer load and finds the optimal value of P. We propose a novel mechanism that enables a single-round MapReduce job to collect statistical information about the input data and the underlying cluster performance, determine and impose a load balancing scheme and optimize the selection of the pivot position for the HII algorithm. The tuning job operates as follows:

- For each possible pivot position P, we estimate a *key distribution*: a mapping between each partitioning key and the number of records that correspond to that key (*i.e.* its cardinality).
- For each pivot position, we devise a greedy *key partitioning* scheme that balances as much as possible the load between reducers;
- For each pivot position, using performance measurements gathered throughout the tuning job execution, we use a cost model to *predict the runtime* associated to each pivot position. We then select the pivot position that yields the shortest job runtime.

Because the search space for an optimal pivot position is small and capped by n – the number of grouping sets – our mechanism can afford to evaluate all values of P. We minimize the overhead of the tuning job by executing it as a single MapReduce job, where the input data is read only once and all candidate values for P are evaluated in parallel. The rest of this section presents how each step of our tuning job is accomplished. Key partitioning described in Section 4.2.3.3.1, uses the estimation of key distribution using sampling (Section 4.2.3.3.2). From partitioning outputs, we extract the reducer loads to feed into our cost model, as shown in Section 4.2.3.3.3. The cost model uses a regression-based approach that predicts the runtime of both mappers and reducers for each value of P; this model uses performance measurements that we collect in the tuning job as a training set.

4.2.3.3.1 Balancing Reducer Load Skewed data motivates the need for reducer load-balancing. In this step, we balance reducer load using cardinalities of partitioning keys obtained from the key distribution estimation discussed in Section 4.2.3.3.2. The input of this step is a set of keys K and an estimation of their respective cardinalities

$C = \{C_k | k \in K\}$. We then perform load balancing by partitioning the keys in K on r reducers to minimize the input keys on the most loaded reducer. This problem can be reduced to the multi-way number partitioning problem, which is NP-complete [70]. We thus opt for a greedy solution: sort all reduce input keys by descending cardinalities, and assign at each step a key to the least loaded reducer. Our algorithm uses smaller cardinalities to counter the imbalance created by large cardinalities that are allocated first. Its runtime is linear with respect to the number of keys. Thanks to the role of combiners, the number of input records sent to each reducer is quite small; for this reason, as we shall see in Section 4.2.5.2, our load balancing strategy is very effective in practice.

The output of this greedy algorithm is a reducer partitioning table, which we broadcast to all mappers to determine the key-reducer mapping. When the Rollup query is running, such a partitioning is kept in a hash table and used to "route" keys to reducers. If, due to sampling, a key does not appear in the hash table, it is "routed" to a pseudo-random reducer using hashing. We note that the impact of such missing keys is minimal on load balancing, as they correspond to infrequent input data.

4.2.3.3.2 Sampling and Computing Key Distribution The load balancing step we just discussed requires the number of reducer input records per grouping key, *i.e.* its cardinality. Without modifying the MapReduce architecture, we cannot count these cardinalities on the fly when processing data. Instead, we resort to *sampling*: we only read data from a small subset of the input data, and we gather key distributions and cardinalities that are the input of the load balancing step. These steps are counted as an overhead in our total job runtime.

In MapReduce, uniform record random sampling from the whole input would be inefficient, as it requires the whole data to be read and parsed. Instead, we employ *chunk random sampling*. Chunk sampling allows low overhead, as it reads a very small portion of input data. The dataset is split in chunks of data that have different sizes: this is also useful for the linear regression model we describe in Section 4.2.3.3.5. Each chunk is chosen with probability σ, a parameter we explore in our experimental evaluation.

However, chunk sampling introduces sampling bias. To reduce this bias, for each chunk, we output each record only once, regardless of its multiple appearances in that chunk. Such a technique can be easily achieved by using combiners. In the reducers, we collect records from different chunks and treat them like records gathered from uniform random sampling. This method, albeit in a different context, is described in the literature as the COLLAPSE approach in [42]. For large input data, the COLLAPSE approach is proved to be approximately as good as uniform random sampling.

For each value of P, we gather statistics from the sample data. The statistics are collected from both mappers and reducers in the tuning job, which are used for two goals: *(i)* to construct the key distribution by examining the histogram of partitioning key; *(ii)* to

gather input of performance measurements in each phase for our cost model. In this way, the work done in this step is also used to benchmark system performance and provide input for cost model. This is an improvement to other standard sampling methods.

4.2.3.3.3 Cost-Based Pivot Selection In Section 4.2.3.2 we have discussed qualitatively the impact of the pivot position over the runtime of map and reduce phases. We now describe our pivot selection technique in detail. Here, we present a cost model to predict the runtime of a Rollup query implementing the HII algorithm, and use it as a reference to optimize the value of P.

4.2.3.3.4 The Model The Rollup query runtime T_J is defined as a function of P:

$$T_J(P) = \overline{T_M}(P) * \alpha + T_{R_{\max}}(P), \tag{4.3}$$

where $\overline{T_M}$ represents the average runtime of a map task, α is the number of map *waves* and $T_{R_{\max}}$ is the runtime of the slowest reduce task.

Waves are due to the fact that the number of map slots can be smaller than the number of input splits to read. The number of waves is computed as

$$\alpha = \left\lceil \frac{\text{input size}}{\text{input split size} \cdot \text{number of map slots}} \right\rceil. \tag{4.4}$$

We assume here that the number of reduce tasks will not be larger than the number of available reduce slots (indeed, such a setting is generally not recommended in MapReduce), therefore for simplicity, we consider the reduce phase to have one wave only.

We decompose the map and reduce phases into several steps. The mappers *read* the input, *replicate* the data, *write* map output records to memory, *sort* output records, *combine* and *spill* output to local disk. For some steps such as *read*, *parse* and *replicate*, their runtime does not depend on the pivot position. Since we are interested in finding the value of P that minimizes the running time rather than predicting the running time itself, we focus on minimizing the variable parts of T_M:

$$\begin{aligned} T'_M(P) = {} & c_{\text{write},P}(2\beta) + c_{\text{sort},P}(2\beta) \\ & + c_{\text{combine},P}(2\beta) + c_{\text{spill},P}(\beta'_P). \end{aligned} \tag{4.5}$$

Here, β is the number of input tuples of a mapper; β' is the number of output tuples of its combiner; $c_{\lambda,P}(x)$ is a cost function that returns the runtime of step λ (write, sort, combine, spill) and depends on a particular value of P. Since HII generates 2 outputs for each input tuple, we have 2β as the input of $c_{\lambda,P}(x)$.

Similarly to the map phase, the reducers *shuffle* and *merge* their input records, *process* the Rollup operator and write outputs to a distributed file system (*DFS*). We therefore estimate the reduce phase as:

$$T_R(P) = c_{\text{shuffle},P}(\gamma_P) + c_{\text{process},P}(\gamma_P) + c_{\text{DFS},P}(\gamma'_P), \tag{4.6}$$

where γ_P is the number of reducer input records and γ'_P is the number of reducer output records.

4.2.3.3.5 Regression-Based Runtime Prediction We propose a novel approach to predict the runtime of each step. This approach is not only completely transparent to users but also flexible enough to deal with any cluster configuration. We achieve such flexibility and transparency by using the tuning job we introduced in Section 4.2.3.3.2 to obtain performance measurements. We apply regression on these measurements using the least squares method to predict the runtime of the Rollup query.

Performance Measurements The tuning job can be thought of as a preliminary Rollup job, *i.e.* it is composed by the same steps of the Rollup query. For each step, we measure the number of its input records and its runtime as a data point (x, T_x). To gain more accurate runtime prediction, we generate several data points. In the map phase, the tuning job runs multiple mappers on small, different chunk sizes; each chunk size is a data point[6]. It also measures the runtime of the *process* phase and *DFS* I/O at several different points in time during the reduce phase.

Linear Regression Let us consider a step λ: we model its cost as a linear function $c_{\lambda,P}(x) = a + bx$. The sort step is an exception because of its $\mathcal{O}(n \log n)$ complexity: $c_{\text{sort},P}(x) = a + bx + cx \log x$. In the tuning phase, we obtain for each step a list of data points $[(x_1, T_{x_1}), (x_2, T_{x_2}), \ldots, (x_\mu, T_{x_\mu})]$; we use least square fitting to find the coefficients a and b such that our function $c(x)$ minimizes the error $S = \sum_{i=1}^{\mu}(T_{x_i} - c(x_i))^2$. This approach necessitates the chunk sizes x_i we use for sampling to be of variable size. Once a and b are known, we calculate the runtime T_λ on the original dataset.

Runtime Prediction and Pivot Selection Now that all $c_{\lambda,P}$ functions are known, we need to evaluate β, β', γ and γ' in order to obtain runtime predictions from Equations 4.5 and 4.6. Again, using regression we can estimate β, β', γ and γ'. We have now all the required values; we can therefore evaluate Equations 4.5 and 4.6 for all possible values of P. The value P^* that minimizes $T'_M(P^*) + T_R(P^*)$ is our chosen value as pivot position.

[6] In our experiments, the chunk sizes are 256KB, 512KB, 1MB, 2MB.

4.2.4 Implementation Details

We now discuss our implementation of the Rollup operator described in Section 4.2.3, for the Apache Pig system. Our implementation is at the base of a patch which has been integrated to the Apache Software Foundation for integration.[7]

Apache Pig is a client-facing system that exposes a high-level language called *Pig Latin*, which resembles standard SQL. End-users express their data analysis tasks in Pig Latin, which Pig transforms into native MapReduce jobs to be executed on an Apache Hadoop cluster. Figure 4.11 illustrates the steps executed by Pig to compile a Pig Latin script into one or more MapReduce jobs [62]. After reading and parsing the script, Pig builds a *logical plan* – an abstract directed-acyclic-graph (DAG) representation of the various operators involved in the script – and proceeds with its optimization. The optimization of the logical plan follows a rule-based approach: the original plan can be rewritten to achieve better performance, *e.g.*, by applying early projections to reduce I/O requirements. Then, Pig transforms the logical plan into a physical plan – a different representation of the script, with physical operators replacing logical ones. Finally, Pig generates and optimizes the MapReduce plan, which contains a set of MapReduce operators that are injected into a Java representation of the original script.

The Rollup Compiler Currently, a data transformation involving a Rollup operator written in Pig Latin has the "`Cube` *alias* `by Rollup` *expression*" syntax. To accommodate our new Rollup operator, we modify the original syntax to expose the sampling rate used in the tuning phase.

In addition, we also enrich the logical plan optimization engine with our optimizer: by traversing the DAG representation of the logical plan, our optimizer detects the presence of a Cube operator, and checks whether the aggregate function to be applied by the operator is algebraic or not. Then, the optimizer replaces the original Rollup logical operator with our new logical operator: this takes into account how, internally, tuples of data are generated according to the pivot position. As a consequence, also the physical plan now includes a new partitioner and a new Rollup physical operator, that implements the HII algorithm.

The Tuning Job The tuning job, described in Section 4.2.3, computes an optimal pivot position and a partitioning table to achieve load balancing. Recall that the tuning job materializes as a preliminary, short, MapReduce job that needs to be executed before the MapReduce job in charge of executing the Rollup operator.

[7] https://issues.apache.org/jira/browse/PIG-4066

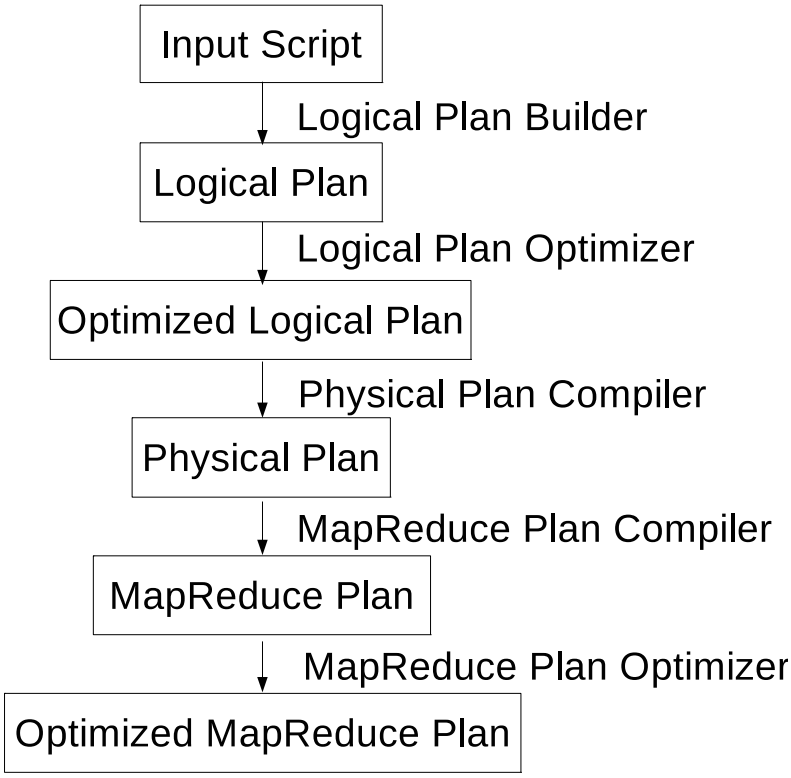

Figure 4.11: Pig script compilation process

As described in Section 4.2.3, we run the sampling of our tuning job on variable-sized chunks. We modify `PigInputFormat` – which specifies how Pig should read input data – to allow such a requirement; the default sizes we use are 256 KB, 512 KB, 1 MB and 2 MB and they are configurable via the `pig.hii.rollup.variablesplit` configuration parameter. We select blocks so that a fraction σ (0.0005 by default) of the whole dataset is sampled, splitting it equally between chunk sizes and ensuring that at least one chunk per configured split size is read.

We also modify the MapReduce plan compiler: we insert a new MapReduce operator representing the tuning job right before the MapReduce HII operator. Hence, our tuning job is executed before the HII Rollup job. In the tuning job, we introduce "markers", that are used to measure the map task runtime, which we also use in the cost-model. Finally, the reduce task of the tuning job, once the pivot position has been set according to the cost-model, outputs a partitioning table that is instrumental for the execution of the Rollup job.

At the end of the tuning job, the Rollup job begins with an initialization phase: the partitioning table is pushed to the Hadoop Distributed Cache, a mechanism which allows to distribute read-only files to Hadoop workers; as such, the partitioning table is replicated across all machines involved in the Rollup job execution. In practice, upon instantiation

of the partitioner class of the Rollup job, the partitioning table is loaded into memory and it is used to "route" map output tuples to the right reducer.

Additional System-level Optimizations In the map phase of Hadoop MapReduce, map tasks collect intermediate output into a memory-buffer, which is divided into two parts: one stores the meta-data of a processed record, the other stores the record itself. If either one of these two parts grows above a threshold, the map task spills both buffers to disk, which is a costly operation involving disk I/O. In the optimal case, we want both buffers to reach the thresholds at the same time to utilize the memory-buffer to its full extent and avoid spilling as much as possible.

In Pig, meta-data always takes up 16 bytes, while the size of map output records varies. Fortunately, we sample the average size of output records $\overline{q}$ during the tuning job: we therefore configure the value of the `io.sort.record.percent` parameter to its optimal value $16/(16 + \overline{q})$.

4.2.5 Experimental Evaluation

We now proceed with an experimental evaluation of our Rollup operator, implemented for Apache Pig. Our experimental evaluation is done on a Hadoop cluster of 20 machines with 2 map and 1 reduce slot each. The HDFS block size is set to 128MB. We execute Rollup aggregates over date-time dimensions using synthetic and real-life datasets; our reference performance metric is *job runtime*, with jobs being executed in an isolated cluster. We note that all results in the following are the average value of at least 10 runs: the standard error of results is smaller than 3%, so for the sake of readability, we omit error bars from our figures.

4.2.5.1 Datasets and Rollup Queries

In our experiments, we use 4 illustrative datasets. Each dataset has tuples with schema *year, month, day, hour, minute, second* and a value v. The Rollup operator computes the total value $V = \sum v$ per each date-time dimension and as a whole (*i.e.*, the `total` level). Specifically, we used the following datasets:

- **Synthetic Telco Logs (STL)**: 1 billion records ranged in 30- years with uniform distribution.
- **Skewed Synthetic Telco Logs (SSTL)**: 1 billion records. The tuples are in a 3-year time-frame, according to a power-law distribution with coefficient $\alpha = 2.5$.

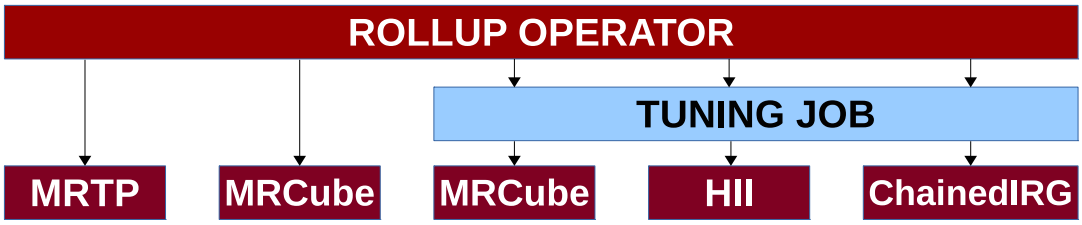

Figure 4.12: Rollup operator with our tuning job

- **Simplified Integrated Surface Database(ISD)**:[8] nearly 2.5 billion records in 114 years. This dataset is heavily skewed towards recent years, as the last 10 years contain 45% of the total records.
- **Reverse DNS (RDNS)**: a sub-set of the Internet Census2012 [9]. It has 3.5 billion records spanning 5 months.

4.2.5.2 Experimental Results

We now present our experimental evaluation that uses a prototype implementation of our Rollup operator for the Apache Pig system. Figure 4.12 illustrates the different flavors of our operator that we used in our experimental campaign.

First, we provide a comparative analysis of our operator using 4 Rollup algorithms: the standard MRTP, and our implementation of MRCube, HII and ChainedIRG. This series of experiments are an illustration of the versatility of our approach, that yields a Rollup operator that can achieve substantial performance gains over current standards.

Next, we focus on the tuning job, customized for the HII Rollup algorithm: we show that our cost-based optimizer can indeed find the most suitable operating point to achieve minimum job runtime. In addition, we measure the overhead imposed by the tuning job, and relate it to its optimization accuracy.

Finally, to validate the efficiency of our design, we compare the overhead of the tuning job customized for MRCube against the original 3-phase MRCube design, and show that even such an algorithm could benefit from our approach.

4.2.5.2.1 Comparative Performance Analysis In this series of experiments, we execute a simple Rollup query, as shown below, on the four datasets described above, and measure the runtime required to complete the job using different flavors of the Rollup operator.

[8] http:// www.ncdc.noaa.gov/oa/climate/isd/
[9] http://internetcensus2012.bitbucket.org/

```
A = LOAD path/file AS (y, M, d, h, m, s, v);
B = CUBE A BY ROLLUP(y, M, d, h, m, s) RATE samplingRateValue;
C = FOREACH B GENERATE group, SUM(cube.v);
```

Figure 4.13 compares the runtime of different Rollup operators using MRTP, MRCube, HII and ChainedIRG algorithms for all datasets. On top of the bars of MRCube, HII, and ChainedIRG, we indicate the the gain of each algorithm with respect to MRTP: for example, -26.83% on top of MRCube means that the corresponding jobs terminate 26.83% quicker than with the current Apache Pig implementation.

For all datasets, our operator for the HII algorithm outperforms the MRTP implementation by at least 50%. Indeed, the map phase of MRTP generates 7 output tuples for each input one, while HII only generates at most 2 tuples. When customized for the ChainedIRG algorithm, our operator runs faster than HII, as its map phase generates only 1 tuple. For RDNS dataset, both HII and ChainedIRG degenerate to $P = 1$ which is the IRG algorithm, which explains the identical runtime.

When compared to MRCube, both the HII and ChainedIRG variants perform better. The first reason is that the mappers in MRCube generate more tuples than HII and ChainedIRG: 3 for ISD and STL, 4 for SSTL and even 5 for RDNS datasets. The second reason is that the reduce phase of MRCube incurs redundant computation. The third reason is that our tuning job runs faster than the first job of MRCube (that would correspond only to a fraction of our tuning job's mission: data sampling and cardinality estimation).

4.2.5.2.2 Cost-based Parameter Selection Validation In this series of experiments, we override the automatic selection of the pivot position, but allow the load balancing at reducers, to proceed with a "brute- force", experimental approach to validate the cost-based optimizer of our tuning job.

For all datasets, we run with the full-fledged tuning job and compare the pivot position output by the cost-based optimizer to that yielding the smallest job runtime, for all possible (manually set) positions. Due to the lack of space, we only present results for the SSTL and ISD datasets, but we obtain similar results for all other datasets. Also our ChainedIRG operator exhibits the same pattern we present here for the HII algorithm. In the following Figures, we report (on top of each histogram) the increment (in seconds) in query runtime for sub-optimal pivot positions.

Query runtime as a function of P for the SSTL and ISD dataset are reported in Figures 4.14 and Figure 4.15, respectively. First, we note that $P = 1$ is essentially IRG. It has the fastest map phase, but the worst performance. This is due to the lack of parallelism which results in longer shuffle and reduce runtime. At the other extreme, $P = 6$ also

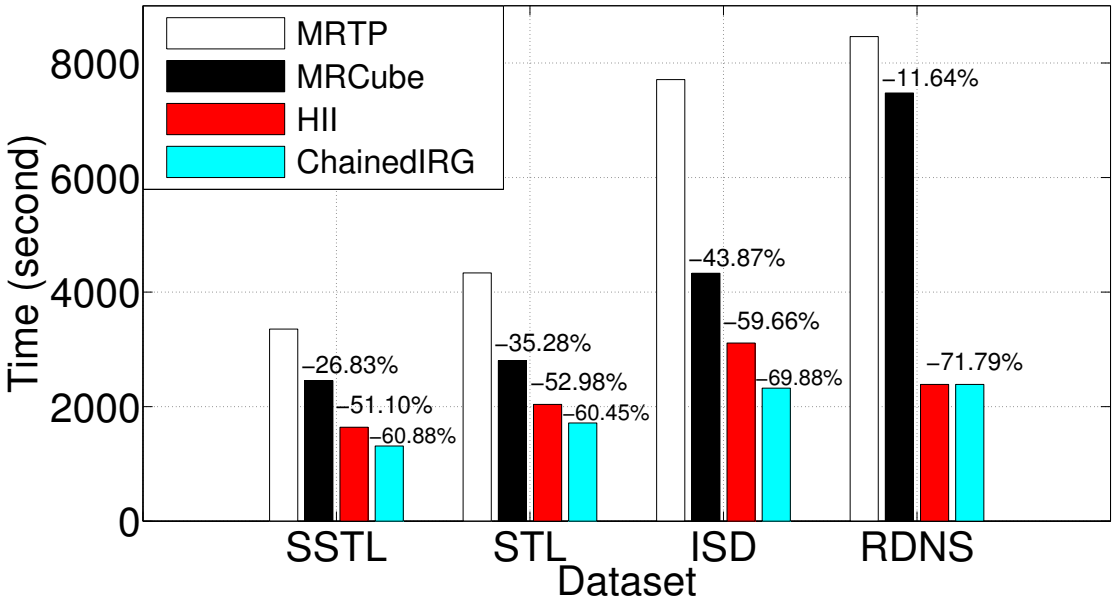

Figure 4.13: Job runtime of four approaches, 4 datasets

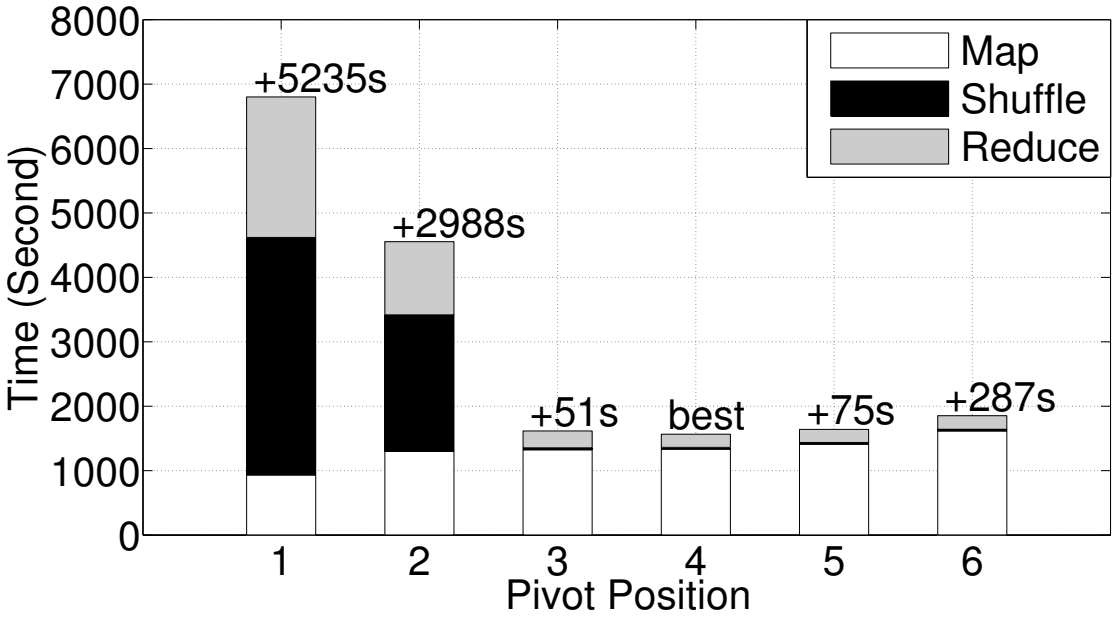

Figure 4.14: Job runtime with Pivot runs from 1 to 6, SSTL dataset

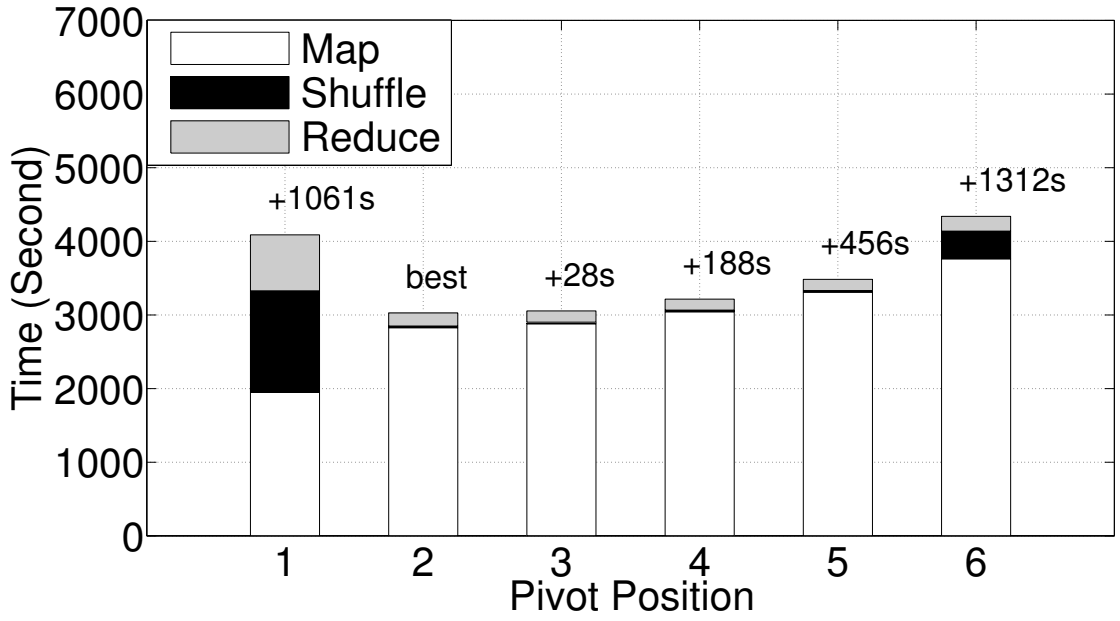

Figure 4.15: Job runtime with Pivot runs from 1 to 6, ISD dataset

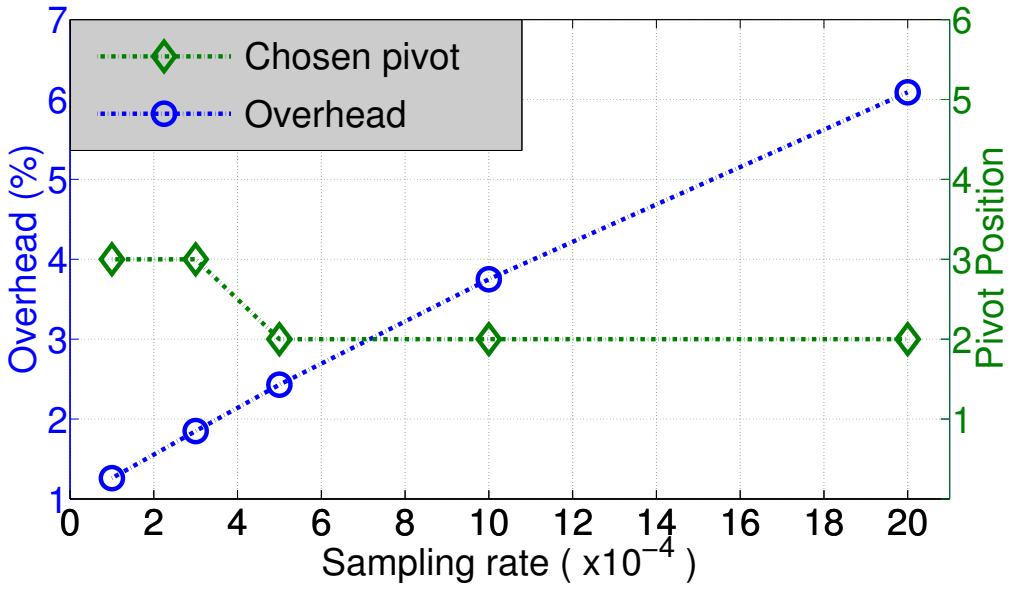

Figure 4.16: Overhead and accuracy trade-off, ISD dataset

results in performance loss for several reasons, including low combiner efficiency (and hence longer shuffle runtime) and longer reducer runtime. The SSTL dataset presents a peculiar case: since the data covers only 3 years, the query runtime corresponding to $P = 2$ can only utilize 3 of the 20 available reducers in our cluster. As a result, both shuffle and reduce phases take a long time to process, due to the lack of parallelism. Instead, in the ISD dataset, the data covers 114 years. It means that $P = 2$ can fully utilize parallelism. Finally, we verified that the pivot chosen in our tuning job correctly specified $P = 4$ and $P = 2$ as the optimal pivot positions for the SSTL and ISD datasets respectively.

4.2.5.2.3 Overhead and Accuracy trade off of the Tuning Job We define the overhead as the runtime of the tuning job divided by the total runtime of the Rollup operator. We examine this overhead as a function of sampling rate. We also study the trade-off between the overhead and the accuracy of the pivot position determined by our cost model. We plot the overhead of the tuning job, along with the pivot position that the cost model selects with sampling rate from 0.0001 to 0.002. Again due to lack of space, we only show the plot for the ISD dataset as a representative result.

In Figure 4.16, due to data skew, a low sampling rate does not produce sufficiently accurate cardinalities, and as a consequence, the cost model and the load-balancing phase under-perform: they cannot determine the optimal pivot position. Nevertheless, the pivot position chosen by our tuning job quickly stabilizes to the optimal value ($P = 2$). As a consequence, the sampling phase only imposes less than 3% overhead of the whole operator runtime.

4.2.5.2.4 Efficiency of Tuning Job Finally, to conclude our experimental evaluation, we verify our tuning job efficiency in comparison to that of MRCube. We integrate load balancing into MRCube and call it MRCube-LB. In MRCube-LB, we replace the first round of the MRCube algorithm by our tuning job: the cost-based optimizer is disabled, while data sampling, cardinality estimation and load balancing are active. Next, we present a notable improvement of the MRCube-LB with respect to the originial MRCube. Figure 4.17 shows, for all datasets, the *overhead* of our operator in three versions: MRCube, MRCube-LB and HII algorithms. Both the overhead of the HII and MRCube-LB versions are always smaller than that of the original MRCube by at least 60%. Comparing the HII and MRCube-LB versions, the marginally higher overhead of the former is due to the execution of the cost-based optimizer.

In conclusion, this series of experiments indicate that our tuning job is versatile, lightweight, and accurate.

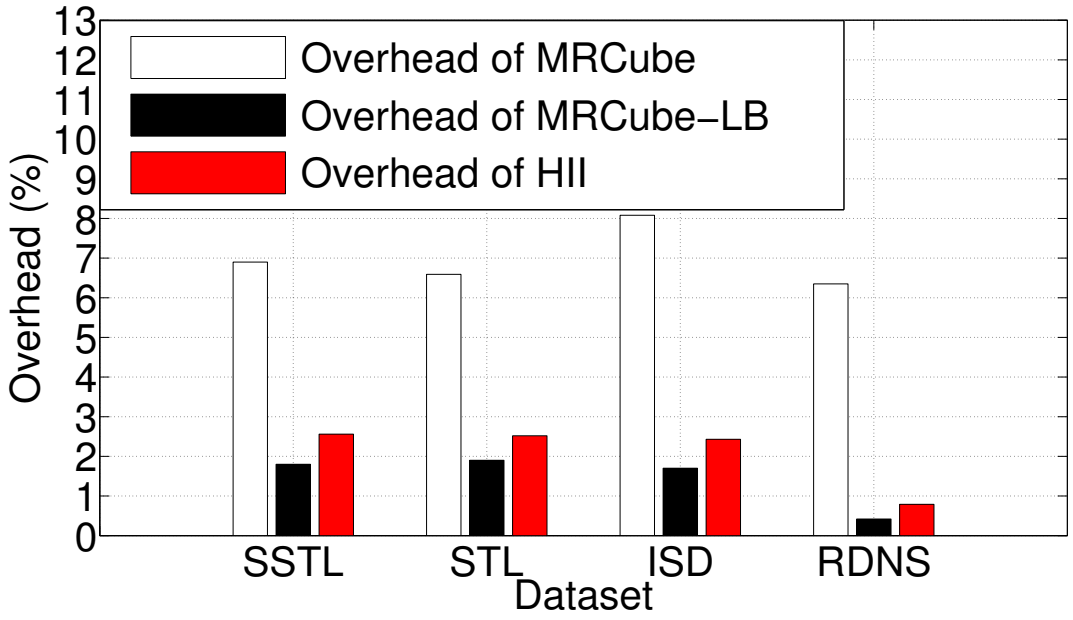

Figure 4.17: Overhead comparison of MRCube, MRCube-LB and HII.

4.3 Rollup as the Building Block for Data Aggregation

We conclude this Chapter by noting that the physical optimization for MapReduce Rollup constitutes as a building block to provide physical optimization for the multiple data aggregation problem as well. Figure 4.18 shows an example of an optimized logical plan obtained from our logical multiple Group By optimization in Chapter 3. Let us consider the following computing order: $T \dashrightarrow ABCD \rightarrow AB \rightarrow A$ which T is the input file, one can see that it is a variant of MapReduce Rollup, without computing ABC and $*$. We note that, actually our definition of Rollup operator in Section 4.1.3 already covers this situation. Therefore, our algorithms as well as our Rollup design can be directly used for the multiple Group By optimization problem.

4.4 Summary

In this Section, we have studied the problem of logical optimization of computing Rollup aggregates in MapReduce. This problem requires:

1. Efficient algorithm(s) for MapReduce Rollup.
2. A physical Rollup operator that is executable using MapReduce.

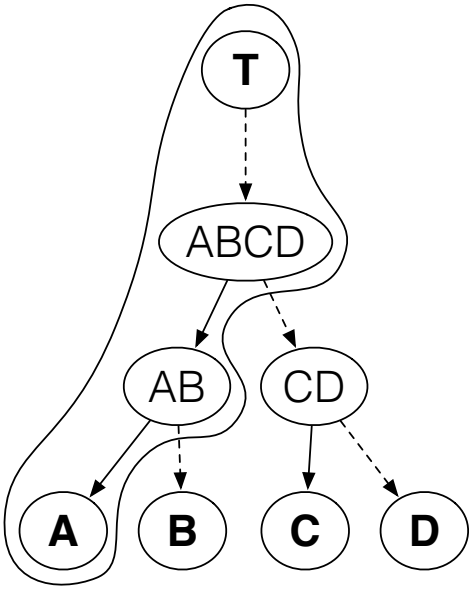

Figure 4.18: An example of optimized logical plan.

To design efficient algorithms, from the theoretical view, we proposed a modeling approach to untangle the available design space of the MapReduce Rollup problem (Section 4.1). We focus on the trade-off that exists between the achievable parallelism and communication costs that characterize the MapReduce programming model. This was helpful in identifying the limitations of current Rollup implementations, that only cover a small portion of the design space as they concentrate solely on parallelism. We presented an algorithm to meet the lower bounds of the communication costs we derived in our model, and showed that minimum replication can be achieved at the expenses of parallelism. Furthermore, we presented several variants of Rollup implementation algorithms that share a common trait: a single parameter (the pivot) allows tuning the parallelism *vs.* communication trade-off.

Our experimental evaluates several Rollup algorithms (including ours and state of the art ones) using Hadoop MapReduce. The experimental approach revealed the importance of optimizations currently available in systems such as Hadoop, which could not be taken into account with a modeling approach alone. Our experiments show that the efficiency of the new algorithms, which we design, is superior to what is available in the current state of the art algorithms. A version of this theoretical work is published in BeyondMapReduce 2014 [28].

Having designed such efficient algorithms, we move on to Section 4.2 to discuss how to employ our algorithms in practice. The main contribution of Section 4.2 was the design of an efficient, skew-resilient Rollup operator for MapReduce high-level languages. Its principal component, the tuning job, is a lightweight mechanism that materializes in a small job executed prior to the Rollup query. The tuning job performs data and performance sampling to achieve, at the same time, cost-based optimization and load balancing of a range of Rollup algorithms (including ours and even state of the art ones).

Our extensive experimental validation illustrated the flexibility of our approach, and showed that – when appropriately tuned – our Rollup algorithms dramatically outperform the one used in current implementations of the Rollup operator for the Apache Pig system. In addition, we showed that the tuning job is lightweight yet accurate: cost-based optimization determines the best parameter settings with small overheads, which are mainly dictated by the data sampling scheme. Our work is available as an open-source project[10] and published in BigDataCongress 2015 [29]. We conclude this Chapter by noting that the physical optimization for MapReduce Rollup constitutes as a building block to provide physical optimization for the general data aggregation optimization problem as well.

The last two Chapters have presented the work done in both phase of query optimization: logical and physical for large-scale data aggregation. Still there is a missing piece to glue these two phases together: an optimization engine. In the next Chapter, Chapter 5, we present our design of a multi-query optimization engine.

Relevant Publications

- D. H. Phan, M. Dell'Amico, and P. Michiardi, "On the design space of MapReduce ROLLUP aggregates," in *Algorithms and Systems for MapReduce and Beyond*, 2014.
- D. H. Phan, Q. N. Hoang-Xuan, M. Dell'Amico,and P. Michiardi, "Effcient and self-balanced Rollup aggregates for large-scale data summarization", in *4th IEEE International Congress on Big Data*, 2015.

[10] https://github.com/bigfootproject/AutoRollup

Chapter 5

Multi-Query Optimization Engine for SparkSQL

In this Chapter, we describe our multi-query optimization engine, called SparkSQL Server. Though our work is based on the Apache Spark system, we reckon that the idea behind our work can be applied comfortably to other MapReduce-like systems.

5.1 Introduction

There are two types of query optimization: single-query optimization and multi-query optimization. As its name suggest, single-query optimization optimizes a single query by deciding, for example, which algorithm to run, configuration to use and optimized values for parameters. An example is the work in Chapter 4: when users issue a Rollup query to compute aggregates over *day*, *month* and *year*, the optimization engine automatically picks the most suitable state of the art algorithms and set the appropriate parameter to obtain the lowest query response time.

On the other hand, multi-query optimization optimizes the execution of a set of multiple queries. Typically, in large organizations, many users share the same data management platform, resulting in a high probability of systems having concurrent queries to be processed. A cross industry study [71] shows that not all data is equal: in fact, some input data is *"hotter"* (*i.e.* get accessed more frequently) than others. Thus, there are high chances of users accessing these "hot" files concurrently. This is also verified by in industrial benchmarks (TPC-H and TPC-DS) in which their queries frequently access the same data. The combined outcome is that optimizing multiple queries over the *same input data* can be significantly beneficial: multi-query optimization is able to cut down the global execution time remarkably compared to executing each query separately.

In traditional databases or data warehouses, a query is engaged in both types of query optimization. There have been hundreds of researches that study this long, well-known subject. Arguably, query optimization is one of the main reasons to which traditional databases (or data warehouses) owes for their tremendous success.

As a consequence, for large-scale systems like Hadoop and Spark, there have been a lot of on-going attempts and effort to fill in this gap. There are independent systems such as Apache Hive for Hadoop, or built-in systems like SparkSQL for Spark let users write their queries in their SQL-alike languages (*e.g.* HiveQL for Hive, standard SQL for SparkSQL). Then, these queries are compiled, optimized and executed using the underlying execution engine (*e.g.* Hadoop, Spark). We note that the query optimization engine provided in these systems only deals with a single query from a single user. Thus, a lot of beneficial multi-query optimization techniques for MapReduce and its extensions cannot be implemented inside these systems.

From all above perspectives, we see that a multi-query optimization is the last stone to completely fill the mentioned gap to provide users with the utmost level of performance. In this Section, we propose a multi-query optimization engine for SparkSQL. Our engine accepts multiple SparkSQL queries, and use many plugged in techniques to optimize them. Because of its generality and extensibility, our engine provides a framework upon which new multi-query optimization techniques can be easily implemented and executed. Even though we choose SparkSQL as our main target, we reckon that the main idea of our design can also be applied to other high-level language systems. To summarize, our main contributions in this Section are:

- We design a multi-query optimization engine for SparkSQL. Certainly there are many different techniques for multi-query optimization, therefore our engine is designed to be able to plug in many different techniques. The engine also provide a general framework upon which users can easily implement other techniques. Such an engine, to the best of our knowledge, is the first one for MapReduce-like systems.
- We implement a prototype of our engine for Apache Spark 1.6 [1], which is fully functional and available as an open source software [2]. We call our prototype *SparkSQLServer*.
- Using our SparkSQLServer prototype, we present a cost model for the multiple data aggregation query optimization in Spark and use it to implement three algorithms, including ours and two other state of the arts. This demonstrates the generality and extensibility of our design to easily facilitate different techniques.

[1] Spark latest stable release version
[2] https://github.com/DistributedSystemsGroup/sparksql-server

- We conduct experiments to evaluate in an end-to-end manner to validate the actual benefit of our logical optimization algorithm, and also to demonstrate the merit of our engine.

The rest of the Section is constructed as follows. In Section 5.2, we briefly introduce Spark and SparkSQL. We then describe the design of our multi-query optimization engine for SparkSQL in Section 5.3. In Section 5.4, we discuss our cost model for the multiple data aggregation query optimization problem in Spark. We also discuss the implementation of our technique as well as others in SparkSQLServer. Section 5.5 is dedicated to our experimental evaluation to compare different techniques. Finally we summarize our future work in Section 5.6.

5.2 Apache Spark and SparkSQL

Nowadays, the complexity of data processing systems has grown fast. As first, Google's MapReduce provided a simple and general model for batch processing with fault-tolerance. But MapReduce is not suitable for many other kinds of workloads, such as machine learning, graph processing and data streaming. Each kind of workloads requires a specialized system and model that can be significantly different from MapReduce. Some examples of such systems are Apache Storm [11], Impala [12], and GraphLab [13]. In this context, Apache Spark was created. Its main goal is to provide a unified programming abstraction that can capture different processing models. In this Section, we present:

- An overview of Apache Spark, which covers the Spark Job Submission process.
- A brief introduction of SparkSQL and its main components.

5.2.1 Apache Spark

Apache Spark is an open source project and an implementation of Resilient Distributed Datasets (RDDs) [17] which is, in its turn, a simple extension of MapReduce. More details about RDDs can be found in Chapter 2. Spark reportedly can run programs up to 100 times faster in comparison to the open source implementation of MapReduce: Apache Hadoop [1]. Also, Spark provides a convenient language-integrated programming interface in the Scala programming language but users can also write applications using Java, Python or R from the APIs it supports. Apache Spark has grown rapidly as it is the most active project of Apache.

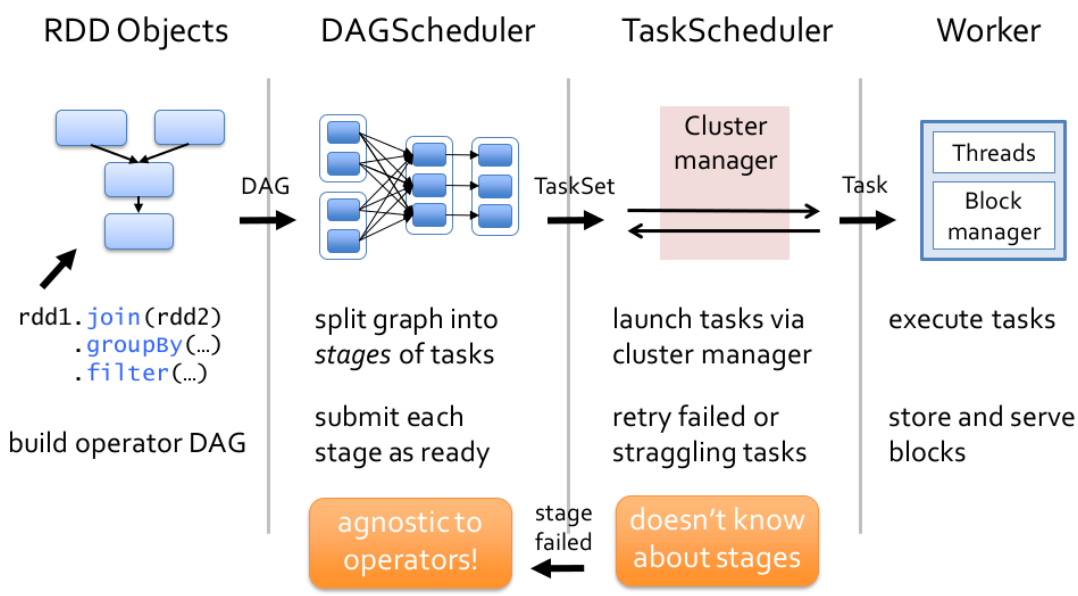

Figure 5.1: The lifetime of a Spark job

5.2.1.1 Spark Job Lifetime

Exploring the Spark job lifetime helps us understanding how a Spark creates and submits a job. This is important since our engine would intercept the Spark job submit procedure: instead of sending Spark jobs directly to the cluster to execute, Spark jobs would be sent to our engines to undergo the multi-query optimization before getting executed.

There are four steps in the lifetime of a Spark Job. Figure 5.1 illustrates those four steps clearly:

1. RDD objects creation.
2. DAG scheduling.
3. Task scheduling.
4. Task execution.

The first three steps happen at the driver, which is our main program. The last step happens at the Executors, where our program is distributed to multiple worker nodes to run in parallel. We go deeper into each step to understand a Spark Job Lifetime.

RDD objects creation: users create RDDs through a set of transformations, at the moment, the RDD is created at the driver.

DAG Scheduling: when an action is called, the DAG which has been built so far is transfered to the next step: DAG Scheduling. It is the process of splitting the DAG into stages, then submitting each stage one after one. All transformations within a stage

are narrow dependencies. Two consecutive stages are connected through a wide dependency transformation. At the boundaries of two stages, a shuffle phase across the whole cluster occurs, and then a module called DAGScheduler, which is in charge of DAG Scheduling, runs these stages in topological order. A stage is a set of independent tasks, which computes the same function as part of a Spark job.

Task Scheduling TaskScheduler is the component which receives the stages from DAGScheduler and submits them to the cluster.

Task Execution Spark calls workers as Executors. It retrieves the worker list from the Cluster Manager, then it launches sets of tasks at the Executor. A BlockManager at each Executor will help it to deal with shuffle data and cached RDDs. New TaskRunner is created at the Executor and it starts the thread pool to process task sets, each task runs on one thread. After finishing the tasks, results are sent back to the driver or saved to disks.

Some information that worth to notice are:

- Thanks to the DAGSchedulerEventProcessLoop, DAGScheduler can keep track of stages' statuses and resubmit failed stages.
- TaskScheduler only deals with task sets after being formed from Stages at DAGScheduler. That is why it knows no stage information.
- A Spark program can contain multiple DAGs, each DAG will have one action. So, inside a driver, they are submitted as jobs one by one.
- Spark has a nice feature: dynamic resource allocation. Spark will base on the workload to request for extra resources when it needs or give the resources back to the cluster if they are no longer used.

5.2.2 Apache SparkSQL

SparkSQL is a Spark module for structured data processing. The interfaces provided by SparkSQL helps Spark know more information about the structure of both the data and the computation being performed. This also gives the flexibility for users to execute relational SQL queries on top of a Spark program. Internally, SparkSQL has a module to use the extra structural information to perform (single) query optimization called Catalyst. Below is a simple example of SparkSQL to demonstrate the combination of relational programming and functional programming of SparkSQL.

```
case class Person(name: String, age: Integer)
val input =
   sc.textFile("examples/src/main/resources/people.txt")
```

```
val people = input.map(_.split(",")).map(p => Person(p(0),
   p(1).trim.toInt))
val peopleDF = people.toDF()
peopleDF.registerTempTable("people")
val teenagers = sqlContext.sql("SELECT name FROM people
   WHERE age >= 13 AND age <= 19")
val result = teenagers.map(t => "Name: " +
   t(0)).collect.foreach(println)
```

First, we declare our data structure, then read the input file with that schema and store into an RDD. We create a DataFrame, a data structure in SparkSQL, from that RDD and create a table from it. We are using the method of reflection to inter-operate between DataFrame and RDD. An SQL query is passed to SparkSQL. After the optimization and translation happened, a normal Spark Job is created with a Directed-Acyclic Graph inside. An action would trigger the job to execute.

5.3 SparkSQL Multi-Query Optimization Engine

In this Section, we present our design of a multi-query optimization engine for SparkSQL that allows many different techniques to be easily implemented and executed.

5.3.1 Multi-query Optimization Techniques

Multi-query optimization techniques can be divided into two categories: *sharing data* and *sharing computation*. Both cases are illustrations of the more generalized *work sharing*: discovering pieces of work in queries that are repetitive and eliminating redundant work. Traditional databases call these pieces of repetitive work *Common Subexpression*.

Sharing data, or scan sharing, means that the common data which are accessible to multiple queries should be scanned and processed only once instead of multiple times. For queries whose significant amount of cost is data scanning, sharing data is crucially beneficial. Most of current MapReduce-like (*e.g.* Hadoop, Spark) multi-query optimization techniques [33, 35, 57] can be casted into this category.

Nevertheless, for many queries (*e.g.* Join, Group By, Cube, Grouping Sets), the major cost are not data scanning but the cost of computation including data shuffling, data sorting, data computing. As a consequence, such *techniques that enable sharing computation in queries for MapReduce-like systems* would be hugely beneficial. As a consequence, our engine should allow both categories of multi-query optimization techniques for SparkSQL.

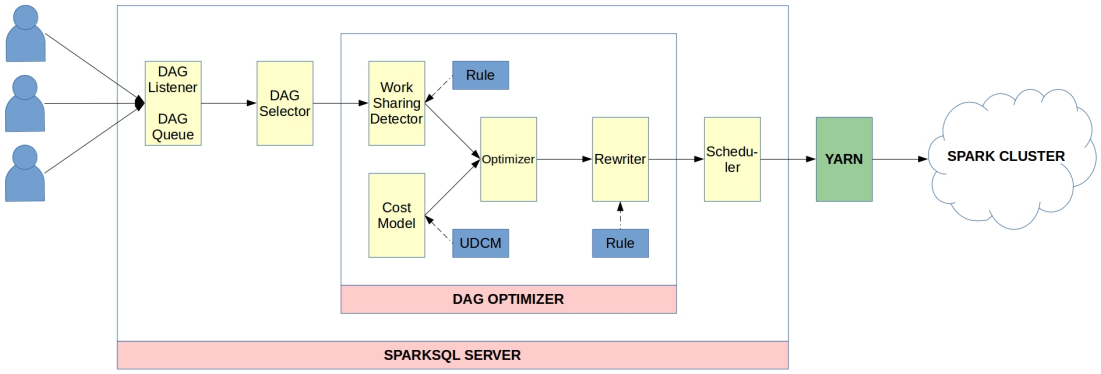

Figure 5.2: The design of SparkSQL Server

Later in Section 5, we present our new technique for sharing computation of multiple Group By queries.

5.3.2 SparkSQL Server

Multi-query optimization works for MapReduce-like systems (hence Spark) currently provide no systematic but rather ad-hoc implementation. Each technique would modify the executing engine (*e.g.* Hadoop, Spark) in their own way, making it hard for users to extend new sharing techniques. To the best of our knowledge, currently there is no unified and systematic framework for multi-query optimization on Spark (and to some extent, other MapReduce-like systems). In this Section, we propose an engine called SparkSQL Server, which serves as a framework upon which other multi-query optimization techniques for SparkSQL can be easily implemented, plugged in and executed. Even though we choose SparkSQL as our main target, we reckon that the main idea of our design can also be applied to other high-level language systems

5.3.2.1 System Design

Our engine aims for the generalization and extensibility so that users can easily implement and plug in other techniques. Figure 5.2 shows our design for such purposes.

Our systems follows the client-server model. **Client side:** clients submit the query and necessary information about data structure, DAG of transformation to SparkSQL Server so that SparkSQL Server can reconstruct the query in the server side

Server side: the SparkSQL Server contains three main components:

- The Listener listens for client connections and communicate with them. It retrieves queries and put them in the DAG queue

- The DAG Optimizer is the heart of our system. All the detections, optimizations, and transformations happen here. The DAG Optimizer consists of:
  - WorkSharing Detector: this module detects the sharing opportunities among a batch of jobs which has just received from clients. The detector uses rule-based mechanism to detect the sharable queries and put them into the same bag. Users can easily write their own rules to detect sharable jobs with many types of sharing.
  - CostModel: it calculates the cost of a plan or a part of a plan. Each technique may have different cost models, and users can provide their own cost model through the User-defined Cost Model interface.
  - Optimizer: it receives the output of WorkSharing Detector, which are a bag of queries that can be optimized. Using the corresponding cost model, the optimizer then optimize this set to return an optimized execution plan.
  - Rewriter: the Rewriter transforms the original execution plan into the optimized execution plan output by the Optimizer and then submits them to the Scheduler. The transformation is also based on the rule-based mechanism. It is very easy to define a new rule to support other techniques.

  Again, all of these four modules are extensible interfaces so users can plug their own implementations into these modules.
- The Scheduler is used when there is a requirement of a specific execution order for queries. This depends on the particular techniques. The Scheduler is also extensible so users can plug their own scheduling strategies.

5.3.2.2 Implementations

Similar to the above subsection, we present the implementation of our system on both client side and server side. The order of the explanation is the same as the flow in figure 5.2.

Client side: we modify the `spark-submit` command so that it sends the application to the SparkSQL Server instead of the cluster manager using this option: `--sparksql-server`

Each client starts its own driver with its own SparkContext. Then, the client generates its DAG from the SQL queries, the DataFrame creation and send those information: DAG, DataFrame generation, SQL Query to the SparkSQL Server.

The client also sends the jar file, because it contains all user defined functions, classes which are very important to reassemble the DAG at the server side. It sends other nec-

essary information so the SparkSQL Server can reconstruct the DAG and the query. We note that the query has already been optimized by SparkSQL at the client side.

Server side: this is the core of our engine

- **DAG Listener and DAG Queue** DAG Listener accepts the client connections and receives DAGs, queries and other information. Then, it puts them to the DAG Queue. The DAG Queue accepts the information the clients send at the FIFO order. In this component, the full DAG of each user will be built based on the initial DAG, the DataFrame creation information and the query. The DAG Queue has a window ω, after reaching the size of the windows, DAG Queue will send a batch of DAGs (and queries also) into the next components.
- **DAG Selector** This component is something called a pre-scheduler, which is based on the constraint attached with each query to fulfill user requirements. For example: a job with the highest priority gets executed right away without further delay of optimization. The default uses the simple FIFO strategy.
- **WorkSharing Detector** This component will detect which DAGs have the opportunity for multi-query optimization. It works as rule-based mechanism to detect the sharing opportunities among queries. The WorkSharing Detector bases on the associated rule to find which DAGs have the Common Subexpression that can be shared. For example, for scan sharing optimization, in which all queries read the same input file, we can build a scan sharing rule. When a RDD reads a file from the disk, it always contains the HadoopRDD with the attached input file path. So our rule tries to find each input file path of each DAG and puts into an array, then it intersects all the arrays to find out which DAG is sharable.
- **Cost Model** It provides an interface so that users can plug their own cost model, which is called User defined cost model, into the engine. We can have many types of Cost Model. The cost model can give score, which is calculated from cost functions, on each DAG or a part of DAG.
- **Optimizer** With a bag from the output of the WorkSharing detector and a cost model associated with its optimization technique, the Optimizer component does the job: build the optimized execution plan. The Optimizer and the CostModel work closely to find out the best result. For example, the Optimizer can generate many combinations of shareable DAGs and give them to the CostModel, the CostModel evaluates them with scores and gives back to the Optimizer. The Optimizer then choose the combination with the best score.
- **Rewriter** This component has many families. Each family will have many rewriting rules. It will generate a rewritten DAG which can be optimized, and use the rule that Optimizer picked to transform the original DAGs into the new ones. Each

sharing technique can have one or more than one rewriter rules. For example, with sharing scan, we can have the rewrite technique using MRShare or caching.

There are some special data structures to wrap a DAG after going through a component:

- *DAGContainer*: use to wrap a DAG when it is received at DAGListener. It contains a DAG, its query and its meta-data that it brings together.
- *AnalysedBag*: use to wrap an array of DAGContainers after it went through the WorkSharing Detector. It contains an Array of sharable DAGContainer. It also contains a label to indicate which type of sharing it is since there are many types of sharing.
- *OptimizedBag*: use to wrap a set of DAGContainers after it goes through the Optimizer, which is the best combination of DAG after going through the Optimizer. It also contains a label to indicate its sharing type and the Rewritten Rule to indicate which rule it will use to rewrite. A no-sharable DAG is also wrapped into an OptimizedBag with the Rewritten Rule equals to Null.
- *RewrittenBag*: use to wrap a rewritten OptimizedBag after the optimized bag went through the Rewriter.

5.3.2.3 An Example on SparkSQL Server

This section introduces an example of SparkSQL Server in action so one can understand how SparkSQL Server works clearly. Three example queries below belong to the scan sharing optimization, where different queries may access the same input file (table). In this example, we employ the MRShare technique [57] to solve such a problem. The detailed implementation of MRShare inside SparkSQL Server is described in our technical report [5].

User1

```
case class Teacher(name: String, age: Integer)
val input = sc.textFile("hdfs://A.txt")
val people = input.map(_.split(",")).map(p => Teacher(p(0),
   p(1).trim.toInt))
val peopleDF = people.toDF()
peopleDF.registerTempTable("people")
val teacher = sqlContext.sql("SELECT name FROM people WHERE
   age <= 50")
val result = teacher.map(t => "Name: " +
   t(0)).collect.foreach(println)
```

User2

```
case class Student(fullname: String, age: Integer)
val input = sc.textFile("hdfs://A.txt")
val people = input.map(_.split(",")).map(p => Student(p(0),
   p(1).trim.toInt))
val peopleDF = people.toDF()
peopleDF.registerTempTable("people")
val student = sqlContext.sql("SELECT name FROM people WHERE
   age >= 19")
val result = student.map(t => "Name: " +
   t(0)).collect.foreach(println)
```

User3

```
case class Person(name: String, age: Integer)
val input = sc.textFile("hdfs://B.txt")
val people = input.map(_.split(",")).map(p => Persion(p(0),
   p(1).trim.toInt))
val peopleDF = people.toDF()
peopleDF.registerTempTable("people")
val adults = sqlContext.sql("SELECT name FROM people WHERE
   age >= 25")
val result = adults.map(t => "Name: " +
   t(0)).collect.foreach(println)
```

Now we analyze the flow of our engine in both client and server sides:

Client Side

- Input: User applications (jar files)
- Output: the information belonged to SparkSQL: JarFile, DAGs, DataFrame creation information and SQL queries are sent to SparkSQL Server. In this examples, we have DAG1, DAG2, DAG3.

Server Side
DAG Listener and DAG Queue

- Input: DAGs and queries from clients
- Output: Array of DAGContainers.
- In the example, let us assume the window size of DAG Queue is 3, so the output will be an array of DAG1, DAG2, DAG3.

DAG Selector

- Input: Array of DAGContainers
- Output: Array of DAGContainers based on scheduling strategies (FIFO at the moment).
- In this simple example, the input and the output is the same as we use FIFO strategy.

WorkSharing Detector

- Input: Array of DAGContainers
- Output: Array of AnalysedBags
- In the example, we got two Bags:
  - DAGBag1: (DAG1, DAG2) with the label: **SCAN-SHARING**.
  - DAGBag2: DAG3 with the label: **NO-SHARING**.

Cost Model

- Input: a combination of DAGContainer which is grouped from MRShare Optimizer
- Output: cost belong to each group. In MRShare, it is a number which is computed through the MRShare's GS function.
- In the example, we use the MRShare Cost Model so it returns a score for each combination. Let us assume the score of group DAG1, DAG2 is the best.

Optimizer

- Input: Array of AnalysedBags
- Output: Array of OptimizedBags after choosing the best score generated by the Cost Model.
- In the example, we got OptimizedBag1: (DAG1, DAG2), OptimizedBag2: (DAG3)

Rewriter

- Input: Array of OptimizedBags
- Output: Array of RewrittenBags
- In the example, since MRShare uses the simultaneous pipeline technique to merge jobs, we got RewrittenBag1: (DAG12), RewrittenBag2: (DAG3)

PostScheduler

- Input: Rewritten Bags of DAGs

- In the example, we got 2 Bags of DAGs. Since the execution order does not affect these jobs or in other way, they are independent on each other, so we just use FIFO strategy to submit to the cluster.

The example is only for scan sharing and just to show the data flow of SparkSQL Server but the system is generalized and extensible so other sharing techniques can be implemented into the system. In the next Section, we introduce another problem of multi-query optimization with many different techniques including ours that are implemented in our SparkSQL Server and are compared to each other in and end-to-end manner.

5.4 Multiple Group By Optimization for SparkSQL

In this Section, we discuss the implementation of three different techniques for Multiple Group By query optimization: the Lattice-Partial Cube (LPC), the Bottom-Up Merge (BUM) and our algorithm, the Top-Down Splitting (TDS) inside the SparkSQL Server engine. By doing so, we are able to:

- Compare the benefit of our algorithm for the multiple Group By query problem in an end-to-end comparison to state of the art algorithms.
- Prove the extensibility and generality of SparkSQL Server that actually allows users to easily incorporate a new multi-query optimization technique.

5.4.1 Cost Model for Spark Systems

We would like to implement our logical algorithm, Top-Down Splitting, using the SparkSQL Server engine to:

- Validate the benefit of our algorithm for the multiple Group By query problem.
- Demonstrate the extensibility and generality of SparkSQL Server to incorporate a new multi-query optimization technique

Thus, we present the cost model for the multiple Group By query problem in Spark as follows:

- Let m be the number of worker nodes.
- Respectively, let $C_r(x), C_w(x)$ be the cost function of reading and writing x data locally: $C_*(x) = ax + b$,
- Let $C_s(x) = axlog_2(x) + bx + c$ be the cost of sorting x data locally

- Let $C_t(x) = ax+b$ be the cost function of transferring x data through the network.
- Let $C_{mem}(x) = ax + b$ be the cost of residing x in memory.
- Let $|u|$ be the cardinality of node u
- Let M_{max} be the maximum amount of data can be resided in Spark's memory. Typically, for large-scale data intensive applications, $M_{max} << |T|$, the input file size.
- Then we have the following cost model:
  - $c_{scan}(u, v) = \begin{cases} C_w(|v|/m) : |v| > M_{max} \mid v \text{ is mandatory,} \\ C_{mem}(|v|/m) : \text{otherwise} \end{cases}$
    Here, a scan cost would occur only when the data is read into memory, thus it incurs no reading cost. If v is a mandatory node, or its size exceeds the memory limit, we have to write it to disks. Otherwise, we store v in memory for future usage.
  - $c_{sort}(u, v) = \begin{cases} C_s(|u|/m) + C_t(|v| * m) : |u| \leq M_{max} \\ C_r(|u|/m) + C_s(|u|/m) + \\ \quad C_t(|v| * m) : (|u| > M_{max}) \mid (u = T) \end{cases}$
    For node $u > M_{max}$, we have to store u on local disks. For $u = T$, u is already on local disks. So in both cases, it incurs a reading cost.

We note that, before SparkSQL Server is up and running, it runs a series of tests to obtain all the cost functions C_r, C_w, C_s, C_t with their appropriate coefficients using linear regression. We note that in MapReduce-like systems, there is a feature called *combiner* to locally aggregate data in the same worker node before sending them through the network. So the size of data going through the network in the shuffle phase will be in the order of node v, not node u.

5.4.2 Implementation of LPC

Below are the following steps that we used to implement LPC inside the SparkSQL Server engine. We note that all these steps can be done even when the SparkSQLServer is up and running.

- Class *MGBDetectorRule* extends the *DetectionRule* class and is registered to *WorkSharingDetector* through the *addDetector* method. The detector, when plugged in with this rule, is able to find and gather all Group By queries from the same input data.

- Class *SparkScanSortGBCost* extends the *CostModel* class and provides all three algorithm the necessary cost functions to operate.
- Class *MGBOptimizerLPC* extends the *Optimizer* class and overrides the *execute* method. The implementation of the *execute* method is the logic of the LPC algorithm. At any step, the LPC algorithm can call *SparkScanSortGBCost*'s method to evaluate its plan.
- Class *MGBRewriteRule* extends the *RewriteRule* class and is registered to *RewriteExecutor* through the *addRewriteRule* method. This rule is essentially states that for any node u that at least 2 out-edges, please cache u before processing its edges.
- Class *MGBScheduler* extends the PostSchedulingStrategy class and is registered to *PostSchedulers* through *addScheduler* method. This class instructs the execution of the solution tree in a depth-first order to minimize memory footprint.

5.4.3 Implementation of BUM and TDS

The implementation of BUM and TDS techniques are very similar to LPC. Typically, different algorithms to solve the same problem can share the detector rule, the cost model, the rewrite rule and the scheduler. In fact, both BUM and TDS implementations reuse MGBDetectorRule, SparkScanSortGBCost, MGBRewriteRule and MGBScheduler. They need to replace only MGBOptimizerLPC with an implementation of their own optimization logic. The two classes we implemented are: MGBOptimizerBUM and MGBOptimizerTDS for BUM and TDS techniques respectively.

5.5 Experimental Evaluation

We now proceed with an experimental evaluation of three different logical optimization algorithms for the multiple Group By query problem. We first implement our multi-query optimization engine, SparkSQL Server. using Spark version *1.6*. We then implement these algorithms using the framework provided by our engine. Our experiments are run on a cluster of 16 nodes, with the HDFS block size set to 256MB.

5.5.1 End-to-End Comparison of Logical Optimization Algorithms

The dataset we use in this experiment is the *lineitem* table from the industrial benchmark TPC-H [23]. It contains 100 million records with 16 attributes. Our query consists of

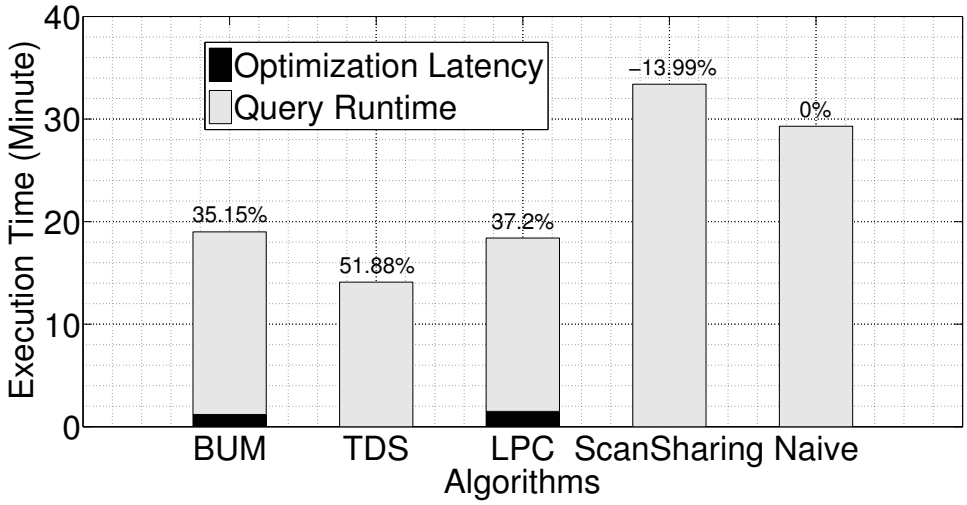

Figure 5.3: The optimization latency and query execution time.

all two-attribute Group Bys from *lineitem*. In our cost model, we use the exact cardinalities obtained from a pre-processing step. For each algorithm, we report, in Figure 5.3, its optimization latency as well as its query runtime. In this figure, we have also the ScanSharing algorithm which is the default implementation of SparkSQL. This algorithm is similar to the *Vanilla algorithm* described in Section 4.1.4.2. It turns out to be the slowest algorithm, even slower than the Naïve algorithm because it produces so much intermediate data and slowdown considerably the sorting and shuffle phase.

Algorithm	BUM	TDS	LPC	SS	Naïve
Optimization Latency (min)	1.2	0	0.15	0	0
Query Runtime (min)	17.8	14.1	16.9	33.4	29.3
Total Runtime (min)	19	14.1	18.4	33.4	29.3
Improvement (%)	35.15	51.88	37.2	-13.99	0

Table 5.1: The optimization latency and query runtime.

Detail numbers found in Table 5.1 (with SS being ScanSharing) confirm that, in large-scale systems, multiple Group By query optimization techniques actually reduce the query runtime over a naïve solution. Again, our algorithm has almost zero optimization latency. Its query runtime is also the lowest among all. This further validates the evaluation that we make in Section 3.5.

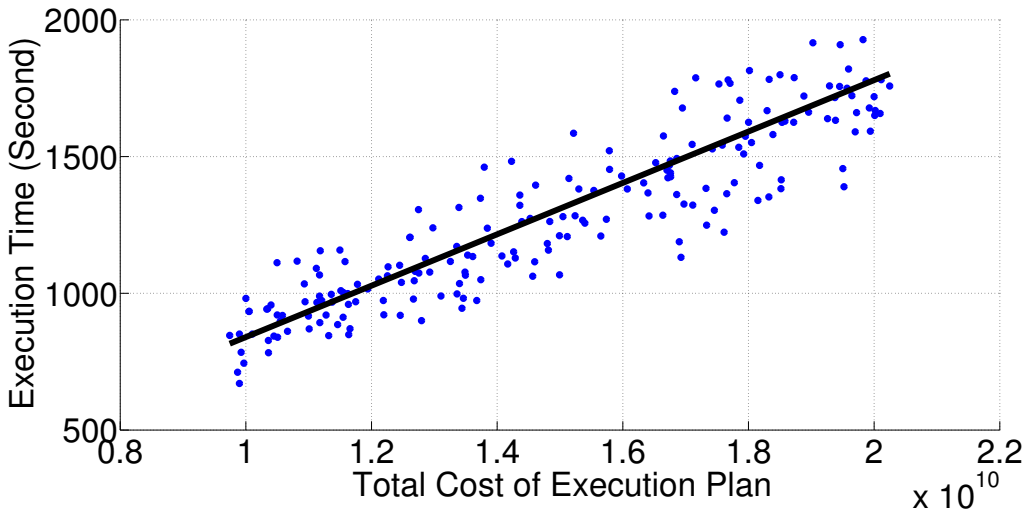

Figure 5.4: Predicted cost vs. actual execution time on our Spark cost model.

5.5.2 The Quality of our Spark Cost Model

The query runtime of the optimized execution plan produced by the logical optimization algorithms depends greatly on the quality of our Spark cost model: they rely heavily on our cost model to guide their optimizing strategy. Ideally, an execution plan of lower total cost should run faster than the one with higher total cost. However, if the quality of our cost model is bad, it may happen that even that our algorithm returns the best execution plan (*i.e.* the lowest total cost), its query runtime is not actually the smallest.

"*Essentially, all models are wrong, but some are useful*" - George E. P. Box. In this experiment, we would like to evaluate the quality of our cost model by generating random execution plans, computing their cost, and then executing them to obtain their query runtime. We plot these plan costs and query runtimes in Figure 5.4. The black line is a linear regression model we built from the plots. It represents an idea of the ideal cost model: any more expensive computation should have higher cost (monotonicity property). As one can see, our model is not "*correct*": there are some higher execution costs with smaller execution times. Still, in general, our cost model pattern still follows the black line, and the differences are within 19%. When compared our cost model to many databases [15], we believe our cost model is performing acceptably well.

5.5.3 Sensitivity Analysis with Cardinality Estimation Error

Our Spark cost model for multiple data aggregation queries depends heavily on the cardinality estimation of grouping nodes. The topic of cardinality estimation is huge, and

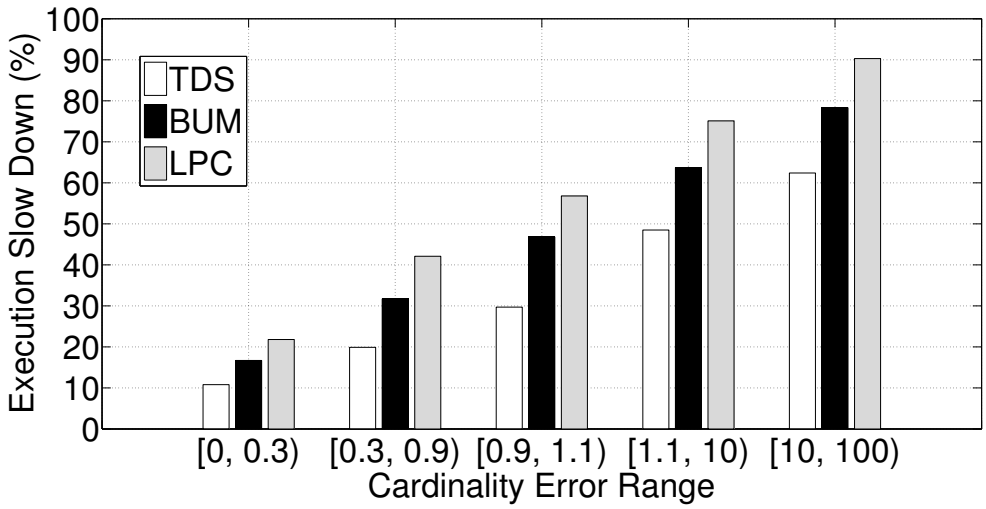

Figure 5.5: The query execution slow down in the presence of cardinality estimation errors.

there has been great effort to improve its correctness and fastness. However, in practice, it is not guaranteed that we can acquire the exact cardinalities all the times. Therefore, in this experiment, we would like to evaluate the sensitivity of these logical optimization algorithms: we would like to see how they behave under the presence of cardinality estimation error.

Using the exact cardinalities of the *lineitem* table obtained from a pre-processing step, we modify these cardinalities by some random factor and calculate their error by the following formula:

$$error(u) = \left| \frac{new\_u - old\_u}{old\_u} \right|$$

Figure 5.5 shows the job execution slow down in the presence of different estimation error. The x axis shows the range of estimation error in form of $[u, v)$, which means that the estimation error of all nodes falls in this range. We see that all algorithms are affected by these errors. LPC seems be the most sensitive to such errors by as much as 90.3%. BUM is less affected and finally, our algorithm TDS is the least affected. We believe that the reason is because when forming an additional node, LPC only considers this node in isolation. Then if the cardinality of this node is estimated greatly off, LPC suffers extensively. On the other hand, BUM merges two nodes and TDS splits multiple nodes at a time, making the effect of a single wrong estimation less influent. The mantra here is that, if two or more cardinalities are off, their correlation to each other may still be right.

5.6 Summary

Multi-query optimization is crucial to provide the maximum performance boost to query processing. There are various types of multi-query optimization, but all should be encompassed in a unified multi-query optimization engine. Such an engine is also the last stepping stone to bridge the gap between traditional databases and modern large-scale systems. This prompts us to design a general and flexible multi-query optimization engine for Apache Spark and SparkSQL.

In this Chapter, we present in details our multi-query engine design that can incorporate different kinds of multi-query optimization. Our engine achieves great flexibility and generality: it can be used as a framework upon which users easily implement their own optimization techniques. We demonstrate these points by implementing several algorithms for the multiple Group By query optimization problem. In addition, scan sharing optimization is also reported in our group work [5]. We then proceed to use our engine to further validate the effectiveness and sensitivity of our algorithm (presented in Chapter 3).

We conclude this Chapter by noting that the idea behind our engine can be comfortably applied to other MapReduce-like systems. We hope that many users will find this engine useful and implement their own techniques to utilize them in production. For researcher on multi-query optimization, our engine should facilitate the design, implementation and evaluation of their work.

Relevant Publications

- This work is currently in preparation to submit to the IEEE Transactions on Knowledge and Data Engineering.

Chapter 6

Conclusion

This dissertation focused on optimizing one of the most predominant operations in data processing: data aggregation for large-scale data-intensive applications. The topic of data aggregation is not new to the database community and there are plenty of works in this domain. Nevertheless, in the context of the current big-data era and large-scale systems like Apache Hadoop or Apache Spark, we have found that current state of the art works are inadequate. They all face the same problem: they run inefficiently and cannot scale to the size that modern data processing tasks require, dealing with data of thousands of attributes and executing thousands of queries.

This dissertation showed that such a problem can only be completely conquered by a thorough combination of optimization algorithm and techniques. While the dissertation provided various contributions in many domains, our main contributions were the logical and physical optimization algorithms and techniques. These optimizations are so intimately related that without one or the other, the data aggregation optimization problem would not be wholly solved. Furthermore, they were integrated as essential components in our multi-query optimization engine that is totally transparent to users. The engine, the logical and the physical optimizations make our works a complete package that was runnable and ready to answer data aggregation queries from users. To the best of our knowledge, this dissertation is the first work to provide comprehensive, efficient and scalable data aggregation for large-scale data-intensive applications using MapReduce-like systems.

Our algorithms and techniques were assessed using both theoretical and experimental approaches. Our theoretical analyses were able to attain the performance and complexity bounds on the worst case and best case scenario, giving our algorithms strong properties on their efficiency. Our experiments were conducted in a real cluster with both synthetic and real-life datasets to further evaluate and enhance our works. Last but

not least, all our works are also available as open source softwares, so that users or other researchers can utilize them for many other purposes.

To conclude, we believe that our works in this dissertation have solved the thesis statement elegantly by a complete solution for efficient and scalable aggregation for large-scale systems from both system and algorithm aspects.

6.1 Future Work

In this Section, using our works as starting points, we present some of the most promising ideas:

Trade-off between multi-query performance and query latency

Multi-query optimization is beneficial when there is redundant work among different queries. The more redundant work queries have, the more performance multi-query optimizations gain. And the more queries we receive in our engine, the more chances of redundant work. This is an incentive for our multi-query optimization engine to wait for as many jobs as possible before performing various optimizations. However, from a user's perspective, this waiting may cause the query to be returned to him/her much later than he/she requires. On the other hand, if our engine executes his/her query right away, it might miss the the chances to find redundant work and drastically improve the query performance. In our engine, this balance is controlled through the DAG queue size ω. Finding the right balance among the queue size, the query latency and the performance gain is a profound but very promising task.

Sharing partially common input

In this dissertation, the condition for multi-query optimizations to happen is when queries access the same common input. This is a correct condition, and has been used in many other works. However, let us consider an extended condition that queries may access the partially common input. For instance, two Group By queries read the same input table, but filter on overlapping predicates (*e.g.* $age < 40$ and $age > 20$). In this case, multi-query optimizations may also spot redundant work, thus reduce the total runtimes. A major future work, which is already going on, is to extend our data aggregation algorithms to consider this case. Moreover, we are developing a more general work sharing technique for multiple queries in Apache Spark that uses its caching mechanism as a primitive.

Unified cost model across different multi-query optimization techniques A query may be subjected to many different types of multi-query optimizations, which leads to an optimization conflict. For example, a Group By query may be optimized using multiple Group By optimization, or scan sharing optimization. The question is:

which optimization applies better to this query. Our engine resolves this conflict using a rule-based approach: when we register the optimization technique, we have to assign a priority level to it as well. If there is a conflict, we choose the highest priority technique. However, sometimes it turns out to be an incorrect decision. A cost-based approach to solve the conflict can be achieved if we have a unified cost model for different techniques. There are some precedents in traditional databases [15], but currently none for large-scale MapReduce-like systems.

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- D. H. Phan, M. Dell'Amico, and P. Michiardi, "On the design space of MapReduce ROLLUP aggregates," in *Algorithms and Systems for MapReduce and Beyond*, 2014.
- D. H. Phan, Q. N. Hoang-Xuan, M. Dell'Amico,and P. Michiardi, "Effcient and self-balanced Rollup aggregates for large-scale data summarization", in *4th IEEE International Congress on Big Data*, 2015.
- D. H. Phan and P. Michiardi, "A novel, low-latency algorithm for multiple group-by query optimization", in 32nd *IEEE International Conference on Data Engineering*, 2016.
- The work in Chapter 5 is currently in preparation to submit to the IEEE Transactions on Knowledge and Data Engineering.

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Appendix A

French Summary

Abstract

Bases de données traditionnelles sont confrontés aux problèmes de scalabilité et d'efficacité avec grandes données. Ainsi, les systèmes de gestion de données modernes qui sont des milliers de machines, comme Apache Hadoop et Spark, ont émergé et sont devenus les principaux outils pour traiter grandes données à grande échelle. Dans ces systèmes, nombreuses d'optimisations de traitement des données qui ont été bien étudiés dans les bases de données sont devenues inutiles en raison des nouvelles architectures et modèles de programmation. Dans ce contexte, cette thèse engagés d'optimiser l'une des opérations les plus prédominants dans le traitement des données: l'agrégation de données pour ces systèmes à grande échelle.

Nos principales contributions étaient les optimisations logiques et physiques pour l'agrégation de données à grande échelle. Ces optimisations sont proches connectés: sans un ou l'autre, le problème d'optimisation d'agrégation de données ne serait pas résolu entièrement. Par ailleurs, nous avons intégré les optimisations dans le moteur d'optimisation multi-requête, ce qui est transparentes pour les usagers. Le moteur, les optimisations logiques et physiques proposées dans cette thèse a formé un package complet qui est exécutable et prêt à répondre aux requêtes d'agrégation des données à grande échelle.

Nous avons évalué nos optimisations théoriquement et expérimentalement. Les analyses théoriques ont démontré que nos

algorithmes et techniques sont beaucoup plus évolutive et efficace que les œuvres antérieures. Les résultats expérimentaux avec un réelle cluster avec des données synthétiques et réelles ont confirmé nos analyses, ont démontré une amélioration significative et ont révélé diverses angles au sujet de nos travaux. En fin, nos œuvres sont publiées comme logiciels libres pour les usages publics.

A.1 Introduction

Les données sont l'atout le plus important pour une entreprise, car ils contiennent des informations précieux. Dans les grandes organisations, les utilisateurs partagent la même plate-forme de gestion de données pour gérer et traiter leurs données, que ce soit une base de données relationnelle, un entrepôt de données traditionnel ou d'un système de grands données moderne. Quelle que soit la technologie sous-jacente, les utilisateurs ou les applications d'analyse de données aimeraient traiter leurs données aussi vite que possible, de sorte qu'ils puissent obtenir rapidement des résultats et prendre des décisions critiques. Encore plus important, l'ère actuelle des grandes données, dans lequel des données a considérablement augmenté en termes de volume et de valeur, a vu des tailles des données de pétaoctets ou même zetaoctets devenir la norme. L'énorme quantité de données met une pression immense sur les systèmes de gestion des données pour traiter efficacement des données de la grande échelle.

Pour répondre à cette demande difficile mais réaliste, les deux critères suivantes doivent être remplies pour permettre traiter efficacement des données à grande échelle:

- Un système de gestion de données qui est capable de passage d'échelle d'un grand cluster (des centaines, des milliers ou plusieurs machines).

- Algorithmes évolutifs et efficaces et des techniques d'optimisation qui sont capables de fonctionner en parallèle sur tous les nœuds du cluster.

Heureusement, le critère premièr a plusieurs réponses adéquates. Actuellement, les systèmes de gestion de données à grande échelle modernes tels que Apache Hadoop [1] ou Spark Apache [2] ont prouvé qu'ils sont capables d'évoluer vers des grappes de mille nœuds [5]. Les entreprises utilisent déjà ces systèmes, ou similaires, pour alimenter leur traitement quotidien des données comme pour calculer trafic web, de visualiser les modèles d'utilisateur, *etc.*. Pour le deuxièm critère, la réponse est de trouver des algorithmes évolutifs et efficaces pour exploiter des puissances de calcul des les systèmes. Ceci est une mission profonde que diverses tâches de traitement de données requiert leurs propres algorithmes et des techniques d'optimisation.

Dans cette thèse, nous nous concentrons de l'une des opérations les plus prédominants dans le traitement des données: l'agrégation de données, ou parfois appelé récapitulation des données. Les utilisateurs qui interagissent avec les grands-données se sentent constamment les besoins du calcul des agrégats pour extraire des idées et obtenir la valeur de leurs données. Bien sûr, les humains ne peuvent pas être attendus pour analyser par téraoctets ou pétaoctets de données. En fait, en général, les utilisateurs interagissent avec les données par des résumés de données. Un résumé des données est obtenu en regroupant des données sur diverses combinaisons de dimensions (par ex-

emple, par emplacement et/ou le temps), et en calculant les agrégats des données (par exemple, compter, somme, moyenne, etc.) sur ces combinaisons. Ces résumés, ou des agrégats de données, sont ensuite utilisées comme données d'entrée pour tous les types de besoins, telles que se joindre à d'autres données, la visualisation des données sur les tableaux de bord, les affaires de prise intelligence décision, l'analyse des données, la détection d'anomalies, etc. Dans cette perspective, nous considérons les données agrégation comme une tâche cruciale qui est effectuée très fréquemment. La charge de travail et d'interrogation des modèles de repères industriels pour les bases de données justifient ce point. Par exemple, 20 sur 22 requêtes dans TPC-H [23] et 80 sur 99 requêtes dans TPC-DS [22] sont des requêtes d'agrégation de données. Cela donne une grande chance pour l'optimisation de l'agrégation des données pour obtenir des performances supérieures.

Cependant, des algorithmes et des techniques d'optimisation disponibles pour l'agrégation des données sur les systèmes à grande échelle modernes sont encore à leurs débuts: ils sont inefficaces et ne pas évolutifs. En outre, en dépit de l'énorme quantité de travail de la communauté de base de données à venir avec e caces façons de calculer les agrégats de données, les architectures parallèles et modèles de programmation distincts de ces systèmes rendent ces travaux incompatibles. En d'autres termes, le problème de l'agrégation efficace des données des systèmes à grande échelle n'a pas les instruments pour répondre à la deuxième critère mentionnée ci-dessus. Ceci est notre moti-

vation complète, et cette thèse est un effort pour pour se remplir le déficit actuel.

Déclaration de la thèse: *Nous proposons et réalisons des nouveaux algorithmes, des techniques d'optimisation et le moteur d'optimisation qui, tous ensemble, fournissent l'agrégation de données évolutive et efficace automatiquement dans les systèmes à grande échelle pour les applications de données intensifs.*

Dans le reste de ce chapitre, nous mettons en évidence nos principales contributions et exposons ce plan de thèse.

A.1.1 Des Contributions et Plan de Thèse

La contribution centrale de cette thèse était une optimisation automatique qui permet l'agrégation de données efficace et évolutive pour les applications gourmandes en données à grande échelle. Ceci est réalisé par deux phases: l'optimisation logique et l'optimisation physique. Toute l'optimisation a été contenu dans un moteur d'optimisation, qui est un élément fondamental de tout système de gestion des données dont l'objet est de trouver le plan de requêtes de la plus haute performance d'exécution.

Les utilisateurs traitent leurs données en émettant des requêtes en utilisant un langage spécifique (par exemple Structured Query Language - SQL) à la plate-forme de gestion des données. Après l'analyse et la validation des requêtes des utilisateurs afin d'assurer qu'ils sont syntaxiquement correct, le système de gestion de données envoie ces requêtes à son moteur

d'optimisation. Ici, pour les requêtes d'agrégation de données, notre moteur d'optimisation effectue d'abord l'optimisation logique en utilisant un de nos algorithmes d'optimisation basés sur les coûts. Puis, il procède à la phase d'optimisation physique, dans lequel on introduit le module d'optimisation poids léger basée sur les coûts. Ce module est capable de: *i)* sélectionner le plus efficace de nos familles de techniques physiques pour se matérialiser des agrégats de données; *ii)* l'équilibrage de la charge de travail entre les différents nœuds d'un cluster pour accélérer les performances. Le résultat de la phase d'optimisation physique est un plan d'exécution optimisé. Enfin, la gestion des données prend ce plan et l'exécute sur le cluster, en utilisant la technique physique sélectionnée de la nôtre. Toutes ces étapes sont effectuées automatiquement et sont totalement transparentes pour les utilisateurs.

Le reste de la présente section est consacrée à un résumé de nos contributions.

A.1.1.1 Optimisation Logique pour l'Agrégation des Données

Notre optimisation commence par la phase d'optimisation logique. Dans cette phase, les données requêtes d'agrégation sont logiquement modélisés à l'aide d'un graphe acyclique orienté (DAG). Parce qu'un DAG est juste un représentant logique du problème, nous pouvons en effet réutiliser de nombreux algorithmes d'optimisation disponibles logiques à partir du domaine de base de données. Cependant, aucun des travaux antérieurs peut évoluer ainsi à un grand nombre de requêtes, ce qui arrive

fréquemment (par exemple dans l'exploration de données ad-hoc), et/ou un grand nombre d'attributs qui sont fréquemment révélées dans des ensembles de données modernes.

Notre principale contribution est de proposer un nouvel algorithme, Top-Down Splitting, qui échelles nettement mieux que l'état des algorithmes d'art. Nous montrons, à la fois théoriquement et expérimentalement, que notre algorithme encourt en très petite l'optimisation des frais généraux, par rapport à d'autres algorithmes, lors de la production des solutions optimisées. Cela signifie que, dans la pratique, l'algorithme peut être appliquée à l'échelle que les tâches modernes de traitement des données nécessaires, traiter les données de centaines ou de milliers d'attributs et l'exécution de plusieurs des milliers de requêtes. Plus encore, cela vient sans aucun sacrifice: en général, notre algorithme est capable de trouver comparable, sinon mieux, que d'autres solutions comme illustré dans notre évaluation expérimentale.

A.1.1.2 Optimisation Physique pour l'Agrégation des Données

La phase d'optimisation physique est en charge de prendre la solution à partir de l'optimisation logique et de décider quelle est la meilleure façon pour la gestion des données système pour le réaliser. La production de cette phase est la stratégie d'exécution définitive qui est plus tard exécuté physiquement sur le cluster. Ainsi, l'optimisation physique dépend fortement du modèle d'architecture et de programmation sous-jacente. Pour le calcul des agrégats de données sur un cluster à grande échelle (par

exemple un cluster Hadoop ou Spark), l'optimisation physique se compose de: *i)* choisir l'algorithme le plus efficace avec les paramètres appropriés; *ii)* d'équilibrage de la charge de travail entre les différents nœuds d'un cluster pour accélérer les performances.

Notre première contribution principale dans cette phase est que nous explorons systématiquement l'espace de conception d'algorithmes pour calculer physiquement agrégats à travers les lentilles d'un modèle de compromis général [49]. Nous utilisons le modèle pour calculer les limites supérieures et inférieures du degré parallèle et le coût de la communication qui sont intrinsèquement présents dans ces grappes à grande échelle. En conséquence, nous concevons et réalisons de nouveaux algorithmes d'agrégation de données qui correspondent à ces limites et le glissement de l'espace de conception, nous avons pu définir. Ces algorithmes se révèlent être remarquablement plus rapide que les œuvres antérieures lorsqu'il est correctement réglé.

Notre deuxième contribution principale est que, nous concevons et réalisons un poids léger, le module d'optimisation basée sur les coûts, ce qui est le cœur de l'optimisation physique. Le module recueille des statistiques de données et les utilise pour choisir correctement l'algorithme et les paramètres les plus efficaces. Ce processus de sélection se fait d'une manière fondée sur les coûts: le module prédit le runtime indicatif de chaque algorithme et les paramètres, puis choisit le meilleur. Last but not least, il effectue également l'équilibrage de charge entre les nœuds du cluster pour accélérer encore la performance.

A.1.1.3 Moteur d'Optimisation de Multi-Requêtes

Aucun des optimisations ci-dessus pourrait fonctionner sans un moteur d'optimisation. Le but de ce moteur est de rassembler et d'orchestrer plusieurs types d'optimisation des requêtes utilisant une requête et de données unifiées représentations. Chaque système de gestion des données a leur moteur d'optimisation propre. Il existe deux types d'optimisation de la requête: une seule requête et multi-requête. L'optimisation simple requête traite chaque requête séparément et indépendamment, tandis que le multi-requête optimise plusieurs requêtes ensemble. Les deux types d'optimisation sont importants, et manque un d'entre eux aurait des résultats dans une performance optimale. En fait, le problème de l'optimisation des agrégations de données comprend à la fois une seule requête et optimisations multi-requêtes.

Nos algorithmes d'optimisation peuvent être mises en œuvre à l'intérieur du moteur d'optimisation des systèmes à grande échelle actuels comme Hadoop et Spark. Cependant, ces systèmes offrent actuellement seuls les moteurs d'optimisation simple requête. Ainsi, notre principale contribution ici est de combler l'écart par la conception et la mise en œuvre d'un moteur d'optimisation multi-requête pour de tels systèmes. Notre conception est flexible et générale qu'il est réellement facile à mettre en œuvre de nombreux types d'optimisation multi-requêtes, y compris la nôtre.

A.1.1.4 Plan de Dissertation

Cette thèse est organisé comme suit. Chapitre 2 présente la fond fondamentale nécessaire pour apprécier pleinement le contexte de cette thèse.

Chapitre 3 est dédié à présenter notre algorithme d'optimisation logique. Nous introduisons notre définition formelle du problème, ainsi que discuter de manière approfondie thestate des algorithmes d'art et leurs limites. Ensuite, nous décrivons en détail notre algorithme et de donner une analyse théorique sur le meilleur des cas et les pires scénarios. Une évaluation expérimentale est menée pour évaluer l'efficacité de l'algorithme par rapport à d'autres travaux.

Chapitre 4 présente nos contributions dans l'optimisation physique pour l'agrégation des données. La première partie de ce chapitre décrit le modèle mathématique que nous utilisons pour calculer les limites des algorithmes d'agrégation de données et de tirer son espace de conception pour correspondre à ces limites. Cette partie se termine par une évaluation expérimentale pour démontrer la compétence de nos algorithmes. La deuxième produ de cette Chapitre est alloué pour le cœur de l'optimisation physique: la conception et la mise en œuvre du module d'optimisation basée sur les coûts. Grâce à nos algorithmes d'agrégation de données, nous montrons sur des ensembles de données synthétiques et réelles que notre design est très léger et a une faible latence d'optimisation.

Chapitre 5 décrit la conception et la mise en œuvre de notre moteur d'optimisation multi-requête. Nous montrons aussi comment mettre en œuvre l'optimisation de l'agrégation des données dans notre moteur, y compris nos techniques ainsi que d'autres œuvres. Les expériences montrent non seulement la flexibilité et de la généralité de notre moteur, mais aussi l'évaluation de bout en bout de différents algorithmes d'optimisation logique pour valider notre travail au chapitre 3.

Le dernier chapitre, le chapitre 6 de cette thèse résume les principaux résultats que nous avons obtenus. Dans la dernière partie de ce chapitre, nous fournissons un ensemble d'orientations futures possibles et discutons notre idée intuitive.

A.2 Optimisation Logique pour l'Agrégation des Données

L'agrégation des données est une tâche cruciale pour comprendre et interagir avec les données. Cette situation est aggravée par la plus grande quantité de données qui sont collectées de nos jours. Ces données sont souvent multi-dimensionnelle, caractérisé par un très grand nombre d'attributs. Cela appelle à la conception de nouveaux algorithmes pour optimiser l'exécution de requêtes d'agrégation de données.

Dans ce chapitre, nous avons présenté notre méthode pour résoudre le problème général de l'optimisation de Groupe multiple par requêtes, comblant ainsi le vide laissé par les propositions actuelles qui ne peuvent pas l'échelle du nombre de requêtes simultanées ou le nombre d'attributs de chaque requête peut gérer. Nous avons montré, théoriquement et expérimentalement, que notre algorithme encourt en très petites latences, par rapport à d'autres algorithmes, lors de la production optimisée des plans de requête. Cela signifie que, dans la pratique, notre algorithme peut être appliqué à l'échelle que les tâches de traitement de données modernes exigent, traiter les données de centaines d'attributs et d'exécuter des milliers de requêtes. En outre, notre évaluation expérimentale illustre l'efficacité de notre algorithme pour trouver des arbres de solutions optimisées. En fait, dans de nombreux cas, notre algorithme a surperformé les autres en termes de production des solutions optimisées, tout en étant remarquablement rapide. Enfin, nous avons discuté à propos de l'intuition derrière notre al-

gorithme et les approches possibles pour l'étendre à la poignée générale, les requêtes hétérogènes en termes de diversité des fonctions d'agrégation. Une version de ce chapitre est publié à International Conference on Data Engineering 2016 [30].

Nous concluons ce chapitre en notant que notre algorithme peut être facilement intégré aux moteurs d'optimisation actuels de bases de données relationnelles, aux entrepôts de données traditionnels ou à des systèmes big-données modernes comme Apache Hadoop [1], Spark Apache [2]. Au lieu de cela, en utilisant notre algorithme pour optimiser l'exécution des requêtes sur des systèmes récents comme Hadoop, Spark et leur haut niveau respectif, interfaces déclaratives, nécessite le développement d'un modèle de coût approprié ainsi qu'un moteur d'optimisation pour transformer les plans de requête d'origine dans les optimisés. Nous couvrons ce chapitre 5. Dans le chapitre suivant, chapitre 4, nous continuons le flux de notre optimisation des requêtes et passer à la phase suivante: l'optimisation physique.

A.3 Optimisation Physique pour l'Agrégation des Données

Dans cette section, nous avons étudié le problème de l'optimisation logique de calcul des agrégats de Rollup dans MapReduce. Ce problème nécessite:

1. Algorithme(s) efficace pour MapReduce Rollup.
2. Un opérateur Rollup physique qui est exécutable en utilisant MapReduce.

Pour concevoir des algorithmes efficaces, du point de vue théorique, nous avons proposé une approche de modélisation pour démêler l'espace de conception disponibles du problème MapReduce Rollup (section 4.1). Nous nous concentrons sur le compromis qui existe entre les coûts de parallélisme et de communication réalisables qui caractérisent le modèle de programmation MapReduce. Cela a été utile pour identifier les limites des implémentations Rollup actuelles, qui ne couvrent qu'une petite partie de l'espace de conception car ils se concentrent uniquement sur le parallélisme. Nous avons présenté un algorithme pour répondre aux limites inférieures des coûts de communication que nous avons tirés de notre modèle, et montré que la réplication minimum peut être réalisé au détriment du parallélisme. En outre, nous avons présenté plusieurs variantes d'algorithmes de mise en œuvre Rollup qui partagent un trait commun: un seul paramètre (le pivot) permet de régler le compromis entre de parallélisme et communication.

Notre expérimentale évalue plusieurs algorithmes de Rollup (y compris la nôtre et de l'état de ceux de l'art) en utilisant Hadoop MapReduce. L'approche expérimentale a révélé l'importance des optimisations actuellement disponibles dans des systèmes tels que Hadoop, qui ne pouvait pas être pris en compte avec une approche de modélisation seul. Nos expériences montrent que l'efficacité des nouveaux algorithmes, que nous concevons, est supérieure à ce qui est disponible dans l'état actuel des algorithmes d'art. Une version de ce travail théorique est publiée dans BeyondMapReduce 2014 [28].

Avoir ces algorithmes efficaces conçus, nous passons à la section 4.2 pour discuter de la façon d'employer nos algorithmes dans la pratique. La principale contribution de la section 4.2 était la conception d'un, biaiser élastique opérateur Rollup efficace pour MapReduce langages de haut niveau. Son composant principal, le travail de mise au point, est un mécanisme léger qui se matérialise dans un petit travail exécuté avant la requête Rollup. Le travail de réglage effectue des données et les performances d'échantillonnage pour atteindre, en même temps, l'optimisation et l'équilibrage de charge basé sur les coûts d'une gamme d'algorithmes de Rollup (y compris la nôtre et même état de ceux de l'art).

Notre vaste validation expérimentale illustré la flexibilité de notre approche, et a montré que - lorsque vous écoutez de manière appropriée - nos algorithmes de Rollup surpassent considérablement celui utilisé dans les implémentations actuelles de l'opérateur Rollup pour le système Apache Pig. En outre,

nous avons montré que le travail de réglage est léger et précis: l'optimisation basée sur les coûts détermine les meilleurs réglages de paramètres avec de petits frais généraux, qui sont principalement dictées par le plan d'échantillonnage des données. Notre travail est disponible en tant que projet open-source [1] et publiée dans BigDataCongress 2015 [29]. Nous concluons ce chapitre en notant que l'optimisation physique pour MapReduce Rollup constitue comme un bloc de construction pour fournir l'optimisation physique pour le problème général d'optimisation de l'agrégation des données aussi bien.

Les deux derniers chapitres ont présenté le travail effectué dans les deux phases de l'optimisation des requêtes: logique et physique pour l'agrégation de données à grande échelle. Pourtant il y a une pièce manquante pour coller ces deux phases ensemble: un moteur d'optimisation. Dans le chapitre suivant, chapitre 5, nous vous présentons notre conception d'un moteur d'optimisation multi-requête.

[1] https://github.com/bigfootproject/AutoRollup

A.4 Moteur d'Optimisation de Multi-Requêtes

Optimisation des multi-requête est cruciale pour fournir le gain de performance maximale pour interroger le traitement. Il existe différents types d'optimisation multi-requêtes, mais tous devraient être englobées dans un moteur d'optimisation multi-requête unifiée. Un tel moteur est aussi le dernier tremplin pour combler le fossé entre les bases de données traditionnelles et les systèmes à grande échelle modernes. Cela nous invite à concevoir un moteur d'optimisation multi-requête générale et flexible pour Apache Spark et SparkSQL.

Dans ce chapitre, nous présentons en détail notre conception du moteur multi-requête qui peut incorporer différents types d'optimisation multi-requête. Notre moteur atteint une grande flexibilité et de généralité: il peut être utilisé comme un cadre sur lequel les utilisateurs facilement mettre en œuvre leurs propres techniques d'optimisation. Nous démontrons ces points en mettant en œuvre plusieurs algorithmes pour le Groupe multiple par problème de l'optimisation des requêtes. En outre, l'optimisation de partage de balayage est également signalé dans notre travail de groupe [5]. Nous procédons ensuite à utiliser notre moteur pour valider davantage l'efficacité et de la sensibilité de notre algorithme (présenté dans le chapitre 3).

Nous concluons ce chapitre en notant que l'idée derrière notre moteur peut être confortablement appliquée à d'autres systèmes de MapReduce-like. Nous espérons que de nombreux utilisa-

teurs trouveront ce moteur utile et mettre en œuvre leurs propres techniques pour les utiliser dans la production. Pour le chercheur sur l'optimisation multi-requête, notre moteur devrait faciliter la conception, la mise en œuvre et l'évaluation de leur travail.

A.5 Conclusion

Cette thèse a concentré sur l'optimisation de l'une des opérations les plus prédominants dans le traitement des données: l'agrégation de données pour les applications de données intensives à grande échelle. Le sujet de l'agrégation des données ne sont pas nouveaux pour la communauté de base de données et il y a beaucoup de travaux dans ce domaine. Néanmoins, dans le contexte des systèmes de l'époque big-données et de grande envergure en cours comme Apache Hadoop ou Apache Spark, nous avons constaté que l'état actuel des travaux d'art sont insuffisantes. Ils sont tous confrontés au même problème: ils courent inefficacement et ne peuvent pas l'échelle à la taille que les tâches modernes de traitement des données nécessaires, traiter les données de milliers d'attributs et d'exécution des milliers de requêtes.

Cette thèse a montré qu'un tel problème ne peut être complètement conquis par une combinaison complète de l'algorithme et des techniques d'optimisation. Alors que la thèse a fourni diverses contributions dans de nombreux domaines, nos principales contributions ont été les algorithmes et les techniques d'optimisation logiques et physiques. Ces optimisations sont si intimement liés que sans un ou l'autre, le problème d'optimisation d'agrégation de données ne serait pas entièrement résolu. En outre, ils ont été intégrés en tant que composants essentiels dans notre moteur d'optimisation multi-requête qui est totalement transparent pour les utilisateurs. Le moteur, le logique et les optimisations physiques font nos œu-

vres un paquet complet qui était runnable et prêt à répondre aux questions d'agrégation des données des utilisateurs. Au meilleur de notre connaissance, cette thèse est la première œuvre pour fournir une agrégation complète, efficace et évolutive des données pour les applications de données intensives à grande échelle en utilisant des systèmes comme MapReduce.

Nos algorithmes et techniques ont été évaluées en utilisant des approches théoriques et expérimentales. Nos analyses théoriques ont été en mesure d'atteindre les performances et la complexité des bornes sur le pire des cas et le meilleur des cas, en donnant à nos algorithmes de fortes propriétés sur leur efficacité. Nos expériences ont été menées dans un véritable cluster avec les deux ensembles de données synthétiques et réelles pour évaluer et renforcer davantage nos œuvres. Last but not least, tous nos travaux sont également disponibles comme des logiciels open source, afin que les utilisateurs ou d'autres chercheurs peuvent les utiliser à d'autres fins.

Pour conclure, nous croyons que nos travaux dans cette thèse ont résolu le rapport de thèse avec élégance par une solution complète pour l'agrégation efficace et évolutive pour les systèmes à grande échelle des deux aspects du système et de l'algorithme.

A.5.1 Travaux Futurs

Dans cette section, en utilisant nos œuvres comme points de départ, nous présentons quelques-unes des idées les plus prometteuses:

Compromis entre Performance Multi-requêtes et la Latence des Requêtes

Optimisation des multi-requête est bénéfique quand il y a travail redondant entre les différentes requêtes. Les requêtes de travail plus redondants ont, plus la performance des optimisations multi-requêtes peuvent gagner. Et plus les requêtes que nous recevons dans notre moteur, plus les chances de travail redondant. Ceci est une incitation pour notre moteur d'optimisation multi-requête d'attendre autant d'emplois que possible avant d'effectuer diverses optimisations. Cependant, du point de vue d'un utilisateur, cette attente peut provoquer la requête à retourner à lui/elle beaucoup plus tard que il / elle a besoin. D'autre part, si notre moteur exécute son/sa requête tout de suite, il pourrait manquer les chances de trouver un travail redondant et d'améliorer considérablement les performances des requêtes. Dans notre moteur, cet équilibre est contrôlé par la file d'attente de DAG taille ω. Trouver le bon équilibre entre la taille de la file d'attente, le temps de latence de la requête et le gain de performance est une tâche profonde mais très prometteur.

Partage d'entrée partiellement commune

Dans cette thèse, la condition d'optimisation multi-requêtes se produire est lorsque les requêtes accèdent à la même entrée

commune. Ceci est une condition correcte et a été utilisé dans de nombreux autres ouvrages. Cependant, considérons une condition étendue que les requêtes peuvent accéder à l'entrée partiellement commune. Par exemple, deux Group By requêtes lire la même table d'entrée, mais le filtre sur le chevauchement des prédicats (par exemple $age < 40$ et $age > 20$). Dans ce cas, les optimisations multi-requêtes peuvent également repérer les tâches redondantes, réduire ainsi le total des runtimes. Un travail futur majeur, qui est déjà en cours, est d'étendre nos algorithmes d'agrégation de données à examiner ce cas. De plus, nous développons une technique de partage du travail plus général pour plusieurs requêtes dans Spark Apache qui utilise son mécanisme de mise en cache comme une primitive.

Modèle de coût unifié à travers différentes techniques d'optimisation multi-requêtes

Une requête peut être soumise à de nombreux types d'optimisations multi-requêtes, ce qui conduit à un conflit d'optimisation. Par exemple, un Group By requête peut être optimisée en utilisant plusiers de Group By optimisation, ou scannez le partage de l'optimisation. La question est: qui applique l'optimisation mieux à cette requête. Notre moteur résout ce conflit en utilisant une approche basée sur des règles: quand nous enregistrons la technique d'optimisation, nous devons attribuer un niveau de priorité à lui aussi bien. S'il y a un conflit, nous choisissons la technique de la plus haute priorité. Cependant, parfois, il se révèle être une mauvaise décision. Une approche basée sur les coûts pour résoudre le conflit peut

être atteint que si nous avons un modèle de coût unifié pour différentes techniques. Il y a des précédents dans les bases de données traditionnelles [15], mais actuellement aucune pour les systèmes à grande échelle comme MapReduce.