Posterior Snapshot Isolation

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Abstract—Snapshot Isolation (SI) is a widely adopted concurrency control mechanism in database systems, which utilizes timestamps to resolve conflicts between transactions. However, centralized allocation of timestamps is a potential bottleneck for parallel transaction management. This bottleneck is becoming increasingly visible with the rapidly growing degree of parallelism of today’s computing platforms. This paper introduces Posterior Snapshot Isolation (PostSI), an SI mechanism that allows transactions to determine their timestamps autonomously, without relying on centralized coordination. As such, PostSI can scale well, rendering it suitable for various multi-core and MPP platforms. Extensive experiments are conducted to demonstrate its advantage over existing approaches.

I. INTRODUCTION

The application of Snapshot Isolation (SI) in real world systems is pervasive. A significant number of the most popular database systems, including Oracle [1], SQL Server [2], PostgreSQL [3], adopt SI as their concurrency control schemes. While SI is weaker than serializability in isolation level, it can avoid blocking caused by read-write conflicts [4]. This gives it tremendous performance gain, especially in applications dominated by read intensive transactions. In recent years, researchers have discovered effective methods to upgrade an SI scheme to support serializability [5]. This further confirmed the suitability and flexibility of SI in database systems.

Snapshot isolation relies heavily on timestamps to determine the temporal relationship among transactions, which allows it to detect conflicts that cause data inconsistency. However, such temporal relationship is, by definition, based on a single clock. To allocate timestamps from a single clock, a central coordinator seems indispensable. As a result, the coordinator can become a potential bottleneck to scalability or a single point of failure for the entire system.

Many of today’s computing platforms come with a high degree of parallelism. On the one hand, due to the weakening of Moore’s Law [6], [7], CPU manufacturers have started adding more and more cores to a single chip to enhance its processing power. If on-chip parallelism keeps increasing, a server with hundreds of cores will be common in the foreseeable future [8]. To prepare for the architectural shift, a number of research projects have recently been launched, aiming to architect database systems on hundreds to thousands of cores [9]. On the other hand, as the data volume in our IT infrastructures grows exponentially, scalability in a large computer cluster is regarded as an important capability of modern database systems. A new generation of parallel database systems, such as those classified as NoSQL [10] or NewSQL [11] databases, have been invented to support such scalability. On a highly parallelized platform, centralized coordination is not desirable, for it may severely impair the scalability and fault tolerance of the system. This prompts us to rethink the design of the snapshot isolation mechanism.

In this paper, we show that SI can actually be implemented without any centralized coordination. The key idea is to delay the assignment of timestamps.

Under SI, if the time intervals of two transactions overlap, they are not allowed to modify the same data. In Figure 1, t1 and t2 conflict, as they both attempt to update Item A simultaneously. This simultaneity can be judged from their timestamps – as the commit time of t1 falls behind the start time of t2, i.e., c1 > s2, their intervals are deemed to overlap.

To protect the consistency of data, we have to abort either t1 or t2. However, this constraint imposed by SI can be overly restrictive. In fact, it is sometimes safe to allow overlapping transactions to modify the same data. Consider the case of t2 and t3, both attempting to modify B. While the intervals of t2 and t3 overlap too, when t3 starts to access B, t2 has already committed. If the two transactions do not share any other data, t3 can safely overwrite the version of B generated by t2, without incurring any inconsistency. In other words, t3 can be regarded as a transaction that starts after the commit of t2. To achieve this, we need to manipulate the timestamps, i.e., c2 and s3, to force c2 < s3.
Instead of using timestamps from a real clock, is it possible to decide the timestamps in the aftermath of transaction execution and induce a logical clock from the timestamps? (An induced clock is shown at the top of Figure[1].) Inspired by this thought, we designed Posterior Snapshot Isolation (PostSI), an alternative scheme of SI that allows transactions to negotiate their timestamps during or after execution. Without relying on a real clock, PostSI manages to get rid of centralized coordination completely. Therefore, it is superior to traditional SI in scalability and fault tolerance (without a central point of failure). In this paper, we introduce the concept of PostSI and a scheduler of it for scalable transaction management over MPP databases. We also introduce a new isolation level called Consistent Visibility (CV), which plays as a stepping stone for realizing PostSI. We conducted extensive evaluation to characterize the performance of PostSI and demonstrate its advantage over alternative approaches.

The rest of the paper is organized as follows. Section 2 reviews the related work. Section 3 introduces the concepts of PostSI and CV as well as the schedulers to enforce their semantics. Section 4 presents the implementation details of PostSI on a shared-nothing architecture. Section 5 reports our experimental results. Section 6 concludes the paper and discusses directions for future work.

II. RELATED WORK

In the literature, implementation of SI on distributed and parallel databases has been active area of research [12], [13], [14]. A number of plausible methods have been proposed, either to improve the scalability of distributed SI [15], [16], [17], [13], [18], or to supplement the SI scheme with data replication [14], [19], [19], [20]. Most of the approaches differ from PostSI in that they still use central clocks to allocate timestamps. In what follows, we briefly review the work that is most relevant to PostSI.

In [13], the authors introduced the concept of Distributed Snapshot Isolation (DSI), an SI scheme for MPP systems. They proposed four methods to implement DSI in a shared-nothing MPP database. Among the four, the pessimistic coordination method and the incremental snapshot method are the most representative ones. In the pessimistic coordination method, each node of a MPP cluster maintains its own clock for timestamp allocation. A local transaction only interacts with the local clock to obtain timestamps. A global transaction needs to block all the participant nodes upfront, to obtain a timestamp from each of the nodes. While this approach does not require centralized coordination, it requires a-priori knowledge about which nodes should be involved in a transaction. Such knowledge is not necessarily available for ad-hoc transactions. Furthermore, initialization of a global transaction is costly in this approach, as it needs to block multiple nodes upfront. By contrast, the incremental snapshot method does not require a-priori knowledge of a transaction, though it requires a central coordinator for managing a global clock. In this approach, when a transaction starts, it only interacts with the local clock of its host. Only when the transaction attempts to access the data on a remote node, does it interact with that node to obtain an appropriate remote timestamp. To ensure the validity of remote timestamps, a mapping between each local clock and the global clock is maintained. Therefore, each node needs to communicate with the coordinator occasionally to keep the mapping up-to-date. Although centralized coordination can be avoided for single-node transactions, it is still mandatory for cross-node transactions. Compared to these approaches, PostSI eliminates the need for centralized coordination completely. Neither does it require any a-priori knowledge about transactions.

To avoid using a central clock, another viable approach is to use synchronized distributed physical clocks (a.k.a. true time devices). A typical example is Spanner [21], a distributed database system developed by Google. Spanner utilizes GPS clocks and atomic clocks to constrain the deviation among different physical clocks within a small error bound. It then builds its concurrency control mechanism upon the timestamps generated by the true time devices. However, as the design of Spanner aims to achieve External Consistency, a stronger level of consistency than SI, its approach is not comparable to PostSI. Moreover, as GPS clocks and atomic clocks are not common hardware, the approach of Spanner does not seem to be universally applicable. Instead of using hardware of high accuracy, Clock-SI [18] resorts to an algorithmic approach that derives timestamps from loosely synchronized physical clocks. Loose clock synchronization [22] would unavoidably result in skew of time. To deal with time skew, Clock-SI has to let a node falling behind to see only old data snapshots or to force an ahead node to wait for a behind node. This makes Clock-SI unstable, as enlarged clock skew will result in severe performance loss. While real clocks can facilitate the synchronization of a distributed database, their maintenance usually incur extra cost. PostSI chooses not to deal with real clocks.

Replication is commonly applied to distributed and parallel databases to enhance their fault tolerance. In [23], Elnikety et al. propose Generalized SI, which allows a transaction to push its start time earlier to facilitate concurrency control on replicated data. In [14], Sovran et al. propose Parallel Snapshot Isolation (PSI), a weaker isolation level than SI that allows different nodes to have different commit orderings. Using asynchronous commit orderings, PSI was shown to achieve significant performance improvement. In [19], an even weaker version of SI called non-monotonic SI was proposed for replicated databases. As non-monotonic SI further relaxes some constraints of PSI, it outperforms PSI in certain circumstances. Other related work on implementing SI over replicated databases can be found in [24], [25], [26], [27], [28], [29], [30], [20]. In this paper, we do not consider data replication. The issue of data replication is actually orthogonal to that of timestamp allocation. Rather than being our competitors, these approaches are complementary to our work.

We are not the first to apply in-transaction or post-transaction timestamp determination. To the best of our knowledge, Timestamp-range Conflict Manager (TCM) proposed by Lomet et al. [31] was one of the earliest attempts at adjusting timestamps during the conduction of transactions, aiming to improve the overall concurrency. TCM was mainly designed for serializable isolation, though it can be applied to weaker isolation, such as read committed, too. It assigns each transaction with a single timestamp to determine its serial order. By adjusting the timestamps on the fly, it allows some
conflicting transactions to proceed concurrently over different versions of the data. TicToc proposed by Yu et al. \cite{yu2015tictoc} applies this idea to avoid centralized coordination in OCC. It marks each data item with two timestamps to represent its valid period. The serial order of a transaction is then decided by synchronizing the time stamps of its read and write sets. In contrast to serializability, SI uses two timestamps to depict the time intervals of transactions (i.e., the beginning and the end of a transaction). While the methods of TCM and TicToc can potentially be extended to work with SI, this feature remains insufficiently explored. Moreover, as they are mainly designed for a single machine database, they do not necessarily work well on a distributed database (see Section 4.1). In contrast, we introduce PostSI in the context of distributed databases.

III. POSTERIOR SNAPSHOT ISOLATION

A. Redefining Snapshot Isolation

Snapshot isolation is a multiversion concurrency control scheme. The conventional definition of SI usually assumes the following two properties \cite{abadi1990snapshot}: (1) Each transaction reads from a consistent snapshot, taken at the start time (also known as snapshot time) of the transaction. A snapshot is consistent if it includes all the writes of the transactions committed before the start time and if it does not include any write of the transactions committed after the start time. (2) Update transactions commit in a total order. In other words, if two concurrent transactions attempt to modify the same piece of data, only one can commit. To determine if two transactions are concurrent, the commit time of each transaction can be recorded; if the time intervals (the interval between the start and commit time) of two transactions overlap, they are regarded as concurrent transactions. While there are several ways to implement SI, all use timestamps as the mechanism to identify the snapshot visible to each transaction.

Instead of using timestamps to determine the snapshots, posterior snapshot isolation first determines the visibility relationship among the transactions. Based on the visibility relationship, it induces the snapshot visible to each transaction and maps the snapshots to a timeline.

Definition 1. (Visibility): Let \( t_i \) and \( t_j \) be two transactions. We say that \( t_i \) is visible to \( t_j \), denoted by \( t_i \rightarrow t_j \), if and only if the writes of \( t_i \) are all accessible to \( t_j \) during the entire lifespan of \( t_j \). We say that \( t_i \) is invisible to \( t_j \), denoted by \( t_i \not\rightarrow t_j \), if and only if none of the writes of \( t_i \) is accessible to \( t_j \) during the entire lifespan of \( t_j \).

In the definition, the writes of a transaction refer to committed writes. We assume that uncommitted writes (intermediate results) are internal data of a transaction and are thus invisible externally. Apart from the visible and invisible relationships, \( t_i \) can be partially visible to \( t_j \), that is, only a subset of the writes of \( t_i \) are visible to \( t_j \). Partial visibility is well possible in a schedule of transactions. However, as it is not acceptable to SI, we exclude it from our discussion. We do not consider temporary visibility either. In other words, a transaction cannot be sometimes visible and sometimes invisible to another transaction. (See Figure 2 for illustration.)

Based on the definition of visibility, the snapshot visible to a transaction \( t \) should comprise: (1) the original version of the database and (2) the writes of all the transactions that are visible to \( t \). If the visibility relationship among all the transactions are determined, the snapshot of each transaction is determined. Then, we know the order of the transactions and which versions of data each transaction should read.

Definition 2. (Visibility Schedule): Given a set of transactions \( T = \{t_0, t_1, t_2, ..., t_n\} \), a visibility schedule of \( T \) is a function \( S : T \times T \Rightarrow \{\text{visible}, \text{invisible}\} \), which maps each pair \( t_i, t_j \in T(i \neq j) \) to either \( t_i \rightarrow t_j \) or \( t_i \not\rightarrow t_j \).

In this paper, we regard a schedule of transactions as a visibility schedule. (For simplicity, we exclude partial visibility and temporary visibility.) In practice, an arbitrary visibility schedule may not be plausible for actual execution. For instance, as illustrated by Schedule I in Figure 3, if two transactions are mutually visible, it is impossible to find an execution plan for the transactions. In order to be executable, a visibility schedule must comply with the data dependencies among transactions.

According to previous work on scheduling \cite{abadi2007visibility}, an actual execution schedule is a flow of interleaved read and write operations of a set of transactions. An execution schedule can introduce three types of data dependency to a pair of transactions:

- \( wr \)-dependency (a.k.a. flow dependency): if \( t_i \) writes data \( A \) before \( t_j \) reads \( A \), \( t_j \) is \( wr \)-dependent on \( t_i \), denoted by \( t_i \rightarrow_{wr} t_j \);

- \( rw \)-dependency (a.k.a. anti-Dependency): if \( t_i \) reads data \( A \) before \( t_j \) writes \( A \), \( t_j \) is \( rw \)-dependent on \( t_i \), denoted by \( t_i \rightarrow_{rw} t_j \);

- \( ww \)-dependency (a.k.a. output dependency): if \( t_i \) writes data \( A \) before \( t_j \) writes \( A \), \( t_j \) is \( ww \)-dependent on \( t_i \), denoted by \( t_i \rightarrow_{ww} t_j \).

Data dependency and visibility relationship are correlated. For instance, if \( t_i \rightarrow_{wr} t_j \), we can conclude that \( t_i \rightarrow t_j \) is false, as \( t_j \) has accessed a write of \( t_i \). Similarly, if \( t_i \rightarrow_{rw} t_j \), we can conclude that \( t_j \rightarrow t_i \) is false. We say that a visibility schedule \( S \) is executable, if there exists an execution schedule \( X \) such that all the data dependencies in \( X \) comply with \( S \). The following is a concrete definition of executable visibility schedules.

![Fig. 2. Cases of Visibility (A and B are the only data shared by T1 and T2.)](image-url)
Definition 3. (Executable Visibility Schedule (or Schedule)): Given a set of transactions \( T = \{t_0, t_1, t_2, ..., t_n\} \), let \( S \) be a visibility schedule of \( T \). We say that \( S \) is executable, if and only if there is an execution schedule \( X \) such that: given a pair of transactions \( t_i, t_j \in T \), (i) if \( t_i \overset{rw}{\rightarrow} t_j \) on \( X \), then \( t_j \rightarrow t_i \) in \( S \); (ii) if \( t_i \overset{wr}{\rightarrow} t_j \) on \( X \), then \( t_i \rightarrow t_j \) in \( S \);

An executable visibility schedule must be compatible with at least one execution schedule. However, an execution schedule may not be compatible with any executable visibility schedule, as it may involve partial visibility, which is not allowed in a visibility schedule. An example is Schedule II in Figure 3. Sometimes, an execution schedule can be compatible with multiple executable visibility schedules – if there is no data dependency between two transactions, their visibility relationship can be arbitrary. In the rest of the paper, we use “schedule” interchangeably with “executable visibility schedule”, which represents both the actual execution of a set of transactions and their visibility relationship.

To ensure that a schedule satisfies SI, we need to map the snapshots induced from the schedule to a linear clock time. Therefore, we define posterior snapshot isolation as follows.

Definition 4. (Posterior Snapshot Isolation): Let \((s, c)\) denote a time interval, where \(s\) and \(c\) are two integers (representing the start time and commit time respectively) and \(s < c\). Let \( I \) denote the domain of time intervals. Given a set of transactions \( T = \{t_0, t_1, t_2, ..., t_n\}\), let \( S \) be a schedule of \( T \). We call \( S \) a snapshot isolated schedule, if and only if we can find a function from \( T \) to \( I \), i.e., \( F : T \Rightarrow I \), that satisfies: (i) given any pair of transactions \( t_i, t_j \in T \) (\( i \neq j \)), either \( t_i \rightarrow t_j \) or \( t_i \rightarrow t_j \) in \( S \) (no partial or temporary visibility); (ii) suppose \( F(t_i) = (s_i, c_i) \) and \( F(t_j) = (s_j, c_j) \), then \( t_i \rightarrow t_j \) in \( S \) if and only if \( c_i \leq s_j \); (iii) if \( t_i \overset{wr}{\rightarrow} t_j \), then \( t_i \rightarrow t_j \).

We say that an execution schedule is snapshot isolated, iff at least one of its compatible visibility schedules is snapshot isolated.

The definition of PostSI requires that the visibility relationship among transactions abides by their temporal relationship – a transaction sees all the transactions that commit prior to its start (Condition (i)). Besides, it also requires that the update transactions commit in a total order (Condition (iii)). Condition (i) requires the visibility relationship between each pair of transactions to be atomic, that is, \( t_i \) is either visible or invisible to \( t_j \). While this condition has been guaranteed by Definition 2, we include it just to make the definition of PostSI complete. In contrast to traditional definition of SI, PostSI does not use a real clock to define the temporal relationship among transactions. Instead, it determines the temporal relationship post priori based on the visibility relationship. As long as there exists a mapping from the visibility relationship to a linear timeline, SI can be satisfied.

For instance, Schedule III in Figure 3 is a PostSI schedule, as we can find appropriate start and commit time for each of its transactions. The timeline at the top of Figure 4 shows the start and commit times that can be induced from their visibility relationship. Based on the physical start and commit time of the transactions, this schedule is not valid to a traditional SI scheduler, which regards \( t_1, t_2 \) and \( t_2, t_3 \) as two pairs of conflicting transactions. However, it is a plausible schedule for PostSI, which uses logical timestamps. In contrast, Schedules IV and V in Figure 3 are not PostSI schedules, as it is impossible to find appropriate start and commit time for their transactions.

If we consider only final effects of scheduling, there is no semantic difference between PostSI and the conventional SI. There have been several variants of SI’s definition in the literature. The Generalized SI defined in \[23\] shares some spirit of our definition of PostSI. It allows a transaction to set an earlier start time than its actual start. In this sense, our definition is even more general or relaxed, as we allow both the start time and commit time to deviate from the actual start and end of a transaction. Time is only logical to PostSI.
B. Consistent Visibility

The objective of our work is to find an scheduler of PostSI that requires no centralized coordination. This can be achieved in two steps. First, we create a scheduler to enforce atomic visibility between each pair of transactions (Condition (i) in the definition of PostSI). Second, we complement the scheduler to ensure that a logical timeline can be inferred from the visibility relationships (Condition (ii) of PostSI). To accomplish the first step, we consider a weaker isolation level called Consistent Visibility (CV), which ensures atomic visibility. CV serves as a stepping stone for us to achieve PostSI.

Definition 5. (Consistent Visibility): Let S be a schedule of a set of transactions $T = \{t_0, t_1, t_2, \ldots, t_n\}$. We say that $S$ satisfies consistent visibility, if and only if it meets the following criteria: (i) given any pair $t_i, t_j \in T (i \neq j)$, either $t_i \rightarrow t_j$ or $t_i \rightarrow t_j$ (no partial or temporary visibility); (ii) if $t_i \xrightarrow{ww} t_j$, then $t_i \rightarrow t_j$.

We say that an execution schedule satisfies CV, iff at least one of its compatible visibility schedules satisfies CV.

As we can see, CV relaxes the requirements of PostSI, by merely abandoning its second condition. Therefore, CV is weaker than SI, for it cannot guarantee a valid temporal order of the transactions. For instance, Schedules IV and V in Figure 3 satisfy the criteria of CV but violate that of SI.

- Schedule IV: Based on the transactions’ operations on $A$ and $B$, we have $t_1 \rightarrow t_2, t_2 \rightarrow t_3$ and $t_1 \rightarrow t_3$. This example shows that the visibility relationship under CV does not satisfy transitivity. However, SI requires visibility to be transitive.

- Schedule V: Based on the transactions’ operations on $A, B, C, D$, we have $t_1 \rightarrow t_2, t_3 \rightarrow t_4, t_3 \rightarrow t_2$, and $t_1 \rightarrow t_4$. To meet the criteria of SI, if $t_1 \rightarrow t_2$, then $t_1$’s logic commit time must be earlier than $t_2$’s logic start time, i.e., $c_1 < s_2$. Analogously, we can deduce $c_3 < s_4$ from $t_3 \rightarrow t_4, s_2 < c_3$ from $t_3 \rightarrow t_2$, and $s_4 < c_1$ from $t_1 \rightarrow t_4$. The inequations $c_1 < s_2, s_2 < c_3, c_3 < s_4$ and $s_4 < c_1$ are cyclic and cannot be all satisfied. Therefore, Schedule V is not a SI schedule.

CV is an isolation level stronger than Read Committed and Repeatable Read (by definition), which suffer from partial or temporary visibility. CV can ensure atomic visibility and a total order of write operations (Condition (ii)). The property of atomic visibility has been mentioned in [34], which proposed a scheduler for the isolation level called Read Atomicity (RA). However, as RA does not guarantee a total order of write operations, it is strictly weaker than CV. As pointed out in Bailis’ thesis [35], RA tolerates the anomalies of Lost Updates and Missing Dependencies defined in Adya’s thesis [36]. In contract, CV does not suffer from these anomalies.

Our methodology is to first create a decentralized CV scheduler, and then build a PostSI scheduler on top of it. According to Definitions 4 and 5, the only additional constraint SI imposes on CV is the assignment of appropriate start and commit time to each transaction. In other words, the sufficient and necessary condition for a CV schedule to be a SI schedule is the following one.

**Theorem 1.** Let $S$ be a CV schedule of a set of transactions $T = \{t_0, t_1, t_2, \ldots, t_n\}$. $S$ is also an SI schedule, iff there is a function from $T$ to the domain of time intervals $I$, i.e., $F : T \Rightarrow I$, such that for any two transactions $t_i$ and $t_j$, their time intervals, $F(t_i) = (s_i, c_i)$ and $F(t_j) = (s_j, c_j)$, meet the following constraints:

- If $t_i \rightarrow t_j$, then $c_i \leq s_j$;
- If $t_i \rightarrow t_j$, then $c_i > s_j$;

**Proof:** As $S$ is a CV schedule, it must satisfy Conditions (i) and (iii) of PostSI. To satisfy Condition (ii), we need to prove that $t_i \rightarrow t_j$ is a necessary and sufficient condition of $c_i \leq s_j$. We already have $t_i \rightarrow t_j \Rightarrow c_i \leq s_j$. We can deduce $c_i \leq s_j \Rightarrow t_i \rightarrow t_j$ from $t_i \rightarrow t_j \Rightarrow c_i > s_j$. Therefore, $S$ satisfies Condition (ii) and is thus an SI schedule.

**Theorem 2.** Suppose $S$ is an SI schedule, and $\prec$ is an order of $T$, such that

- $t_i \prec t_j$, iff $t_i \rightarrow t_j$;
- $t_j \preceq t_i$, iff $t_i \rightarrow t_j$;

$S$ is an SI schedule, if and only if: if $\prec$ contains a cycle, then the cycle must contain two consecutive edges of invisibility in the form of $t_i \rightarrow t_k \rightarrow t_j$ (or $t_j \preceq t_k \preceq t_i$).

**Proof:** First, we prove “if”. Suppose $S$ is an SI schedule and $\prec$ contains a cycle. If we assume that there is no case of consecutive invisibility, then each $t_k \rightarrow t_i$ in the cycle must be adjacent to a visibility edge $t_k \rightarrow t_j$. Then, the cycle can be chopped into pieces, where each piece is either a single visibility edge $t_i \rightarrow t_j$ ($t_i \prec t_j$) or a composite edge $t_i \dashrightarrow t_k \rightarrow t_j$ ($t_i \prec t_k \prec t_j$). $t_i \rightarrow t_j$ implies $s_i < c_j, t_i \rightarrow t_k \rightarrow t_j$ implies $s_k < c_k$. These inequations cannot be cyclic. This contradicts our assumption. Therefore, if there is a cycle, it must contain two consecutive edges of invisibility.

Next, we prove “only if”. If $\prec$ does not contain a cycle, $S$ must be an SI schedule, as it will be trivial to find a start time and a commit time for each transaction such that the conditions of Theorem 1 are satisfied. Suppose $\prec$ contains a cycle and the cycle contains two consecutive invisibility edges $t_i \rightarrow t_k \rightarrow t_j$. If we remove $t_k$, the rest of the edges will be acyclic. Then, it will be trivial to assign start and commit times to the transactions that are compatible with the conditions of Theorem 1. If $t_k$ and $t_i \rightarrow t_k \rightarrow t_j$ are considered, given any pairs of $s_i, c_i$ and $s_j, c_j$ (as long as $s_i < c_i$ and $s_j < c_j$), it is always possible to find a pair of $s_k, c_k$ such that $s_k \leq c_i$ and $s_j \leq c_k$. Therefore, as long as the cycle contains $t_i \rightarrow t_k \rightarrow t_j$, we can always find appropriate start and commit times for the transactions so that SI is satisfied.

To conclude, the necessary and sufficient condition of that $S$ is an SI schedule is: if $\prec$ contains a cycle, then the cycle must contain two consecutive edges of invisibility.

Theorem 2 states that a CV schedule is an SI schedule, if its order $\prec$ is either acyclic, or each cycle in $\prec$ comprises two consecutive invisibility relationships. As illustrated by
Schedules IV and V in Figure 3, the schedules do not satisfy SI, just because their \( \prec \) orders are cyclic (i.e., \( t_1 \prec t_2 \prec t_3 \equiv t_1 \) and \( t_1 \prec t_2 \prec t_3 \equiv t_2 \equiv t_1 \)), and the cycles do not contain consecutive invisibility. The proof of Theorem 2 can be analogized to that of the theories behind Serializable SI [5].

The following theorem depicts the gap between CV and Serializability.

**Theorem 3.** Let \( S \) be a CV schedule of a set of transactions \( T = \{t_0, t_1, t_2, ..., t_n\} \). \( S \) is serializable, if and only if there exists a visibility relationship among \( T \) such that:

- For each pair of transactions \( t_i \) and \( t_j \), if \( t_i \rightarrow t_j \), then \( t_j \rightarrow t_i \);
- The \( \rightarrow \) relationship is acyclic;

**Proof:** According to Definitions 3 and 5, the visibility relationship and the data dependencies in a CV schedule satisfy the following rules: (i) if \( t_i \xrightarrow{rw} t_j \), then \( t_j \rightarrow t_i \); (ii) if \( t_i \xrightarrow{wr} t_j \), then \( t_i \rightarrow t_j \); (iii) if \( t_i \xrightarrow{wa} t_j \), then \( t_j \rightarrow t_i \). The necessary and sufficient condition of serializability is that the \( rw \), \( wr \) and \( wa \) dependencies are acyclic. Thus, if \( S \) is serializable, it is possible to find a visibility relationship such that the order \( \prec \) is acyclic. An acyclic \( \prec \) implies that the \( \rightarrow \) relationship is acyclic too. As \( \prec \) is acyclic, there cannot be two transactions \( t_i \) and \( t_j \) satisfying \( t_i \rightarrow t_j \) and \( t_j \rightarrow t_i \) simultaneously. Therefore, \( t_i \rightarrow t_j \) implies \( t_j \rightarrow t_i \).

Serializability requires a total order of transactions, such that each transaction is visible to the transactions behind it. In contrast, CV and SI allow two transactions to be mutually invisible.

In what follows, we introduce a scheduler for CV and then extend it to PostSI. To ensure that a scheduler satisfies CV or PostSI, we need to guarantee that its implementation only allows the visibility relationship defined in CV or PostSI.

**C. A Scheduler for CV**

As the definition of CV does not involve the concept of clock, it is relatively easy to come up with a CV scheduler that needs no centralized coordination. The main concern of CV is atomic visibility. According to Definition 3, the main task of a CV scheduler is to ensure that the data dependencies among transactions does not violate atomic visibility.

First, if there is a \( rw \) dependency between \( t_i \) and \( t_j \), i.e., \( t_i \xrightarrow{rw} t_j \), we must ensure that \( t_j \) is invisible to \( t_i \), i.e., \( t_j \rightarrow t_i \) (Definition 3 (i)). In other words, if \( t_j \) overwrites a data version \( t_i \) has read, \( t_j \) must be invisible to \( t_i \) and \( t_i \) must not read any data generated by \( t_j \). This can be achieved by keeping track of all the \( rw \) dependencies among the ongoing transactions – whenever \( t_i \) attempts to read a data version generated by \( t_j \), it first checks whether \( t_i \xrightarrow{rw} t_j \) exists: if \( t_i \xrightarrow{rw} t_j \) does not exist, \( t_i \) proceeds to read the data; if it exists, \( t_i \) turns to try an older version of the data.

Second, if there is a \( wr \) dependency between \( t_i \) and \( t_j \), i.e., \( t_i \xrightarrow{wr} t_j \), we must ensure \( t_i \rightarrow t_j \) (Definition 3 (ii)). In other words, if \( t_j \) reads a data version generated by \( t_i \), then \( t_i \) is visible to \( t_j \) and all the other data generated by \( t_i \) should be visible to \( t_j \) too. This can be violated, if we let a transaction always read the freshest versions of the data. For instance, if \( t_i \) commits during the execution of \( t_j \), it is possible that \( t_j \) happens to read an old data version overwritten by \( t_i \) and a new data version created by \( t_i \). However, if the system can keep track of the \( rw \) dependencies, this case can be avoided – when \( t_i \) commits, it should check if it has overwritten any data version read by \( t_j \); if it has, \( t_j \xrightarrow{rw} t_i \) is recorded; then, \( t_j \) is forbidden to read any data generated by \( t_i \).

As we can see, atomic visibility can be achieved by tracing the anti-dependency among the ongoing transactions. Thus, our CV scheduler works as follows.

**A CV Scheduler:**

(1) Each transaction is assigned a unique TID. (To generate unique TIDs in parallel, each session can create a TID as a concatenation of its session id and an id from its local counter.)

(2) Each version of a data item is associated with a TID, recording the transaction that created this version. Each version also corresponds to a visitor list, which records the TIDs of the ongoing transactions that has read this version.

(3) The scheduler maintains an anti-dependency table that records the \( rw \)-dependency between the ongoing transactions. Each entry in the table is in the form of \( t_j \xrightarrow{rw} t_i \).

(4) When a transaction \( t_j \) reads a data item, it always reads the latest version that is visible. It starts with the latest version. If the version’s TID is \( t_i \) and it can find \( t_j \xrightarrow{rw} t_i \) in the anti-dependency table, \( t_j \) regards this version invisible (because \( t_j \xrightarrow{rw} t_i \) implies \( t_i \rightarrow t_j \)). Then, \( t_j \) has to try the second latest version and so on. Once \( t_j \) finds the latest visible version, it performs the read, and during the read operation it adds its TID to the visitor list of the version.

(5) When a transaction \( t_j \) writes to a data item, it places a write lock on the item. It unlocks the item only when it commits. Upon the commit, the data versions created by \( t_j \) is immediately visible to other transactions. After \( t_j \) obtains the write lock, it checks if the following criteria can be satisfied: (i) if \( t_j \) has read the data item, the version it has read must be the newest version; (ii) if the newest version’s TID is \( t_i \), then \( t_j \xrightarrow{rw} t_i \) must not be in the anti-dependency table. If one of the two criteria is violated, \( t_j \) has to abort, because one of its concurrent transactions has updated the version.

(6) When a transaction \( t_j \) commits, for every TID in the visitor list of each data version it has updated (let the TID be \( t_i \)), it adds \( t_i \xrightarrow{rw} t_j \) to the anti-dependency table. After the commit, we safely remove \( t_j \) from all the visitor lists, and every \( t_j \xrightarrow{rw} t_k \) from the anti-dependency table.

The examples in Figure 5 illustrate how the CV scheduler executes transactions. The instruments of visitor lists and the anti-dependency table provide sufficient information to determine the visibility relationship between transactions. We can thus ensure that a transaction is either visible or invisible to another transaction. Rule (5) disallows two concurrent transaction to modify the same piece of data. Thus, the scheduler ensures a total order of the write operations. According to Definition 5, this scheduler can enforce CV. Most
importantly, the CV scheduler can be completely decentralized. It does not require any central data structure for coordination, as the visitor lists and the anti-dependency table can all be distributed.

Despite the fact that CV is weaker than SI and serializability in isolation, it can be a practical solution to applications that do not require strong data consistency. However, as CV is not the focus of this paper, we schedule the elaboration of CV’s use cases in our future work.

D. A Scheduler for PostSI

Our scheduler of PostSI is built on top of the CV scheduler. According to Theorem 1, if a scheduler can assign an appropriate time interval to each transaction, it can ensure that the resulting schedule is snapshot isolated. To avoid centralized coordination, our SI scheduler does not rely on a central clock to determine the time intervals. It leaves to the centralized coordination, our SI scheduler does not rely on that the resulting schedule is snapshot isolated. To avoid the focus of this paper, we schedule the elaboration of CV’s use cases in our future work.

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"Complementary Rules of A PostSI Scheduler:

1. When a transaction $t_j$ starts, the lower and upper bounds of its start time and the lower bound of its commit time are initialized as $s_j = 0$, $\overline{s}_j = +\infty$ and $c_j = 0$.

2. Each version of a data item is associated with two timestamps, CID and SID. CID records the commit time of the transaction that created this version. The SID records the maximum start time of the transactions that have read this version.

3. If a transaction $t_j$ reads or overwrites a data version, then the transaction that created the version should be visible to $t_j$. If the data version’s CID is $cid$, $t_j$ updates the lower bounds of its start and commit time, by setting $s_j = max(s_j, cid)$ and $c_j = max(c_j, cid)$.

4. At anytime during the execution of a transaction $t_j$, if $s_j > s_j$, $t_j$ has to abort, as it is no longer possible to find a valid start time for $t_j$.

5. When a transaction $t_j$ commits, it performs the following actions:

a) Determining its own time interval: $t_j$ sets its start time to $s_j$. Let $S$ be the SIDs of the data versions $t_j$ has read. We set $c_j = max\{c_j \cup S\}$. For every $t_i$ such that $t_i \xrightarrow{rw} t_j$ can be found in the anti-dependency table, we set $c_j = max\{c_j, s_i\}$. Finally, we set the commit time of $t_j$ to $c_j = max\{c_j, s_j\} + 1$.

b) Adjusting the orders of its conflicting transactions: For every ongoing $t_k$ such that $t_j \xrightarrow{rw} t_k$, we set $c_k = max\{c_k, s_j + 1\}$ (as $t_j \xrightarrow{rw} t_k$ implies $t_k \xrightarrow{rw} t_j$ and $c_k > s_j$). For every ongoing $t_i$ such that $t_i \xrightarrow{rw} t_j$, we set $\overline{s}_i = min(\{\overline{s}_i, c_j - 1\})$ (as $t_i \xrightarrow{rw} t_j$ implies $t_j \xrightarrow{rw} t_i$ and $c_j > s_i$).

"
lower bound of \( t_i \)'s commit time to be greater than the start time of \( t_j \), depending on whether \( t_i \) or \( t_j \) commits first. This is enforced by Rule (4). The anti-dependency table records only the \( rw \) dependencies of the ongoing transactions. If \( t_j \) commits before \( t_j \) starts, \( t_j \xrightarrow{rw} t_i \) (i.e., \( t_i \rightarrow t_j \)) will not be found in the anti-dependency table. In this case, the scheduler passes \( t_j \)'s start time to \( t_i \) using SID\$s. This is the only mission of SID\$s.

When a transaction finalizes its time interval (Rule (4)(a)), it needs to ensure that the upper and lower bounds are not violated, i.e., \( s_j \leq s_j \leq s_j \) and \( c_j \leq c_j \). Deciding the start time \( s_j \) is simple, as \( s_j \) will not interfere with the future transactions – our scheduler simply sets \( s_j = s_j \). Deciding the commit time \( c_j \) requires more thought. On the one hand, \( c_j \) may lower the \( s \) of the transactions that regard \( t_j \) invisible (Rule (4)(b)). On the other hand, it will also be used as the CIDs of the updated data (Rule (4)(c)), which will in turn raise the \( s \) of the future transactions reading the data. If \( c_j \) is too small or too large, it may cause other transactions to abort. To minimize the chance of abort, our scheduler sets \( c_j \) to the smallest value that is larger than the \( s \) of the transactions regarding \( t_j \) invisible (as stated in Rule (4)(a)).

If we apply the PostSI scheduler to Schedules IV and V in Figure 3, they will not be allowed to pass. \( t_3 \) of Schedule IV will not be allowed to read the newest version of \( B \), as it will make \( s_2 \) greater than \( s_2 \). Similarly, neither \( t_4 \) nor \( t_2 \) of Schedule V is allowed to proceed when attempting to read the newest versions of \( A \) and \( B \).

According to Rule (4), when \( t_j \) commits, for every \( t_j \xrightarrow{rw} t_j \) and \( t_j \xrightarrow{rw} t_k \), it should inform \( t_i \) and \( t_k \) about its time interval. This is where the concurrent transactions conduct negotiation. However, \( t_i \) or \( t_k \) may fail to receive the message before the message arrives. Nevertheless, when \( t_i \) and \( t_k \) commit, they will initiate their orders to \( t_j \). Thus, we can guarantee that the message from at least one direction will arrive safely. The negotiation is guaranteed to take place.

### IV. THE IMPLEMENTATION

Compared to a traditional SI scheduler, our PostSI scheduler introduces additional overheads to concurrency control. While these overheads can be regarded as the price for eliminating centralized coordination, they can be minimized by careful engineering. In what follows, we discuss how to implement PostSI in an MPP database system with a shared-nothing architecture. To process transactions on a shared-nothing architecture, each transaction is allocated to a single computing node, known as the host of the transaction. The host will, in turn, distribute the work to other nodes, which work concurrently to accomplish the tasks of the transaction. Finally, the host employs a commit protocol, such as two-phase commit, to finish the transaction. Our implementation aims to minimize the cost incurred by blocking and communication.

#### A. Distribution of the Operational Data

To implement a distributed PostSI scheduler, a critical issue is the management of the operational data, i.e., the visitor lists, the anti-dependency table and the bounds of each transaction’s time interval. As the operational data is shared among different transactions, cross-node communication cannot be completely avoided. As such, we need to minimize the communication cost.

Data items and their visitor lists can certainly be collocated, as they are always accessed together. Then, no extra cross-node communication will be incurred. As we do not require visitor lists to be persistent, they can be detached from the data and stored in the memory. This can make the maintenance of visitor lists efficient. It is unlikely that all the data are accessed concurrently. Therefore, the space consumption of visitor lists is usually much smaller than that of the data.

For each \( t_i \xrightarrow{rw} t_j \) in the anti-dependency table, we store it on both the node hosting \( t_i \) and the node hosting \( t_j \). Hence, insertion and deletion of the anti-dependencies requires cross-node communication, while lookup of the anti-dependency table can be performed locally most of the time. We implement each anti-dependency table as a hash table, to facilitate its lookup.

The bounds of the time interval of each transaction is maintained by the host of the transaction. As stated in the PostSI scheduler (Rule (3)), the lower bound of the start time (i.e., \( s \)) needs to be updated upon each data access. When a transaction needs to process the data on a remote node, it delegates the work to the remote node. It sends its TID and a copy of its \( s \) to the remote node too, which can update the \( s \) locally while processing the data. After the remote node finishes its work, it sends its local \( s \) along with the results back to the host, which can update the global \( s \). When a transaction is about to commit, it needs to update the interval bounds of the conflicting transactions (Rule (4)(b) of the PostSI scheduler). This may incur cross-node communication.

#### B. Optimizing Read Intensive Transactions

A major advantage of SI lies in that it eliminates the blocking caused by read-write conflicts. This is especially beneficial to read intensive transactions, which may involve the execution of OLAP queries. PostSI is supposed to preserve this advantage of SI. Fortunately, read operations in PostSI are indeed nonblocking. Although each read operation implies an insertion in a visitor list, such an insertion can be completed in an atomic step, without involving locking. Beside nonblocking read, other tactics need to be applied to ensure the efficiency of read intensive queries.

First, when a transaction performs a read, the CV scheduler requires it to lookup the anti-dependency table to confirm the data’s visibility. If the read occurs on a remote read, the lookup will incur cross-node communication. This can be costly, especially for an OLAP-style query that needs to access a large amount of remote data. Fortunately, PostSI does not need to lookup the anti-dependency table when reading data. Instead of using the anti-dependency table, a transaction can use CID to determine visibility – a data item is visible, only if its CID is smaller than the upper bound of the transaction’s start time.

Second, when performing parallel query processing, a transaction distributes the lower bound of its start time (i.e., \( s \)) to the remote nodes, so that they could update the \( s \) locally.
At the same time, the upper bound of its start time (i.e., $\tau$) on the host can possibly be updated by a conflicting transaction (Rule (4)(b) of PostSI). When the transaction receives the $s_s$ from the remote nodes, it may find that $s_s > \tau$ and have to abort. The same abort can occur repeatedly, if the transaction happens to read a hot remote item that is frequently updated. As a remedy, when a transaction aborts, we can retry the transaction by fixing its initial $\tau$ and $s_s$ at the highest CID the transaction encounters before its previous abort. During the retry, the transaction can avoid accessing the data whose CID is higher than $\tau$. Then, the same abort will not be repeated.

Third, when a transaction ends, it is supposed to update the SIDs of the data it has read and remove its TID from all the visitor lists. For a read intensive transaction, this can be costly. On the one hand, it requires a transaction to maintain a huge read set to memorize all the data items it has read. On the other hand, it requires a transaction to perform a large amount of work in the commit phase, which may impair the performance of distributed transactions severely.

As a workaround, we can apply lazy deletion to the visitor lists. Namely, when a transaction ends, it does not update the visitor lists immediately. Only when the next transaction accesses a visitor list will it remove the outdated TIDs from the list. TIDs are usually generated by incrementing a set of TID counters. The current values of the TID counters can be periodically broadcasted to the nodes, to enable the detection of outdated TIDs.

SID is meant to inform a transaction about the max start time of the committed transactions regarding it invisible, so that the transaction can set its own commit time correctly. If the deletion of the visitor lists is delayed, the update of SIDs can be delayed too, as the visitor lists reserve the information about invisibility. Specifically, when removing a transaction's TID from a visitor list, we check the start time of the transaction, and use the start time to update the SID of the data. This requires each computer node to keep the time interval of a committed transaction in its cache for a certain period.

With the optimization, a transaction does not need to maintain a read set. Neither does it need to update a large number of visitor lists and SIDs when committing. As a result, its commit phase can be significantly shortened. This gives PostSI an advantage over some OCC styled approaches, such as TicToc [32], which may have to perform intensive validation work in the commit phase, during which the data in the write set has to be locked.

C. The Commit Phase

Similar to some OCC approaches [37], PostSI does not need to perform real write before the commit phase. Instead, each transaction can keep its write set private. Only when the transaction is about to commit, it applies its write set to the real data. This provides several benefits. First, the period of each write lock can be shortened, so as to enable higher concurrency. Second, deadlock can be avoided, as the write locks can be added in a strict order.

As shown in the PostSI scheduler, three rounds of communication need to be conducted in the commit phase. First, to determine the commit time of a transaction, the lower bounds of the start times of its conflicting transactions need to be retrieved (Rule (4)(a)). Second, after the commit time is determined, the transaction needs to contact the conflicting transactions again to update their bounds of time intervals; at the same time, the anti-dependency table needs to be updated. Finally, the SIDs and CIDs of the data need to be updated. To work with the two-phase commit protocol (2PC), the first round of communication can be integrated into the prepare phase of 2PC, and the last round of communication can be integrated into the commit phase of 2PC. The second round of communication is an additional overhead to PostSI. Fortunately, this extra round of communication occurs only when there is contention. For transactions that do not compete intensively, this cost can be quite marginal.

There is a short interval after a transaction commits and before it makes its updated version available to other transactions. If another transaction happens to read its updated version during the interval, a race condition may arise. In our implementation, we maintain a writer list for each data item being updated. When a transaction, say $t_1$, updates a data item, it adds its TID to the writer list of the item. Only after $t_1$ commits, it removes its TID from the writer list. When another transaction $t_2$ reads the data item, it will first check the writer list. If $t_2$ finds $t_1$ in the writer list, it will regards $t_1$’s version invisible (by enforcing $s_2 < c_2$ in PostSI). This allows us to prevent the race condition during the commit phase.

V. EXPERIMENTAL EVALUATION

In this section, we evaluate the performance of PostSI. We implemented our PostSI and CV schedulers on a distributed in-memory key-value store, which we created for experiment’s purpose [4]. The KV-store supports point queries, simple range queries (using secondary hash indexes), as well as data manipulation operations such as insertion, deletion and update. When deployed on a cluster, the KV-Store partitions a database based on the key values, and assigns one partition to each node. Each node can host multiple sessions for processing user requests and multiple workers for processing transactions. A master node is responsible for maintaining the database, while it never participates in transaction processing.

A. Experimental Setup

Our experiments were conducted on a cluster of 30 virtual machines. Each virtual machine is equipped with two 4-core Intel Xeon E5620 2.4GHz processors and 32GB of DRAM. Each core has a private 32KB L1 cache and a private 256KB L2 cache. At the physical level, every 4 cores share a 12MB L3 cache. The hypervisor of the virtual machines is the OpenStack released by Icehouse in April 2014. The operating systems running on the virtual machines are all CentOS 6.5. InfiniBand is used for networking. According our test, the bandwidth for end-to-end communication is around 1Gbps.

For comparison’s purpose, we implemented several additional SI schedulers in our systems:

1. We implemented a conventional SI scheduler, by adopting the implementation of PostgreSQL 9.4. The SI scheduler of PostgreSQL uses a single timestamp and a snapshot

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1. The source code is available at [https://github.com/PosteriorSI](https://github.com/PosteriorSI)
of ongoing transactions to determine the time interval of each transaction. The timestamp represents the start time of the transaction. The snapshot contains the TIDs of all the ongoing transactions at the start time. A central coordinator, residing on the master, is responsible for allocating timestamps and maintaining a snapshot of the ongoing TIDs. When a transaction starts, it obtains a timestamp and a snapshot from the coordinator. When a transaction terminates, it contacts with the coordinator again to remove its TID from the snapshot.

2) We implemented a **optimal scheduler**, aiming to find out the best possible performance of SI. This scheduler follows the exact procedure of the conventional SI scheduler, except that it requires no centralized coordination. To eliminate the coordination, we simply assign an arbitrary timestamp and an empty snapshot to each transaction. Therefore, the optimal scheduler does not ensure correctness. It is intended to represent an upper bound of SI’s performance.

3) We also implemented **DSI** (the incremental snapshot method) [13] and **Clock-SI** [18], which represents the best existing schedulers to enforce SI in distributed database systems. DSI allows local transactions to be handled by local nodes, while it requires centralized coordination for global/distributed transactions. Clock-SI does not need centralized coordination, but utilizes synchronized physical clocks to determine the time intervals of transactions. These two methods have been discussed in the related work.

To evaluate the performance, we used the benchmark of TPC-C and SmallBank. To store a database in our KV-store, we use a separate key-value hash structure to store each table. Each tuple is treated as a key-value pair – the primary key of a tuple is treated as the key and the remaining part as the value. To support non-primary key queries, which are frequently invoked in TPC-C, we use secondary hash indexes. To conduct experiments on the aforementioned cluster with 30 virtual machines, we used one machine as the master and the others as slaves. The master does not participate in transaction processing. But it works as the central coordinator for SI and DSI, and the central point for clock synchronization for Clock-SI. Unless otherwise mentioned, each slave hosts 8 worker threads, which are fully devoted to transaction processing. For the tests on TPC-C, we installed 5 warehouses on each node. For the tests on SmallBank, we set the scale factor per node to 1 million customers. Transactions are divided into local transactions and distributed transactions. Each local transaction accesses only local data. Each distributed transaction accesses data from 2-3 randomly selected nodes.

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Fig. 6. Clock-SI’s Performance Drops as Time Skew Increases (based on experiment of TPC-C with 8 nodes and 20% distributed txns)

B. **Effects of Time Skew on Clock-SI**

Clock-SI is completely decentralized, though it requires the physical clocks on the slave nodes to be somehow synchronized. We tested Clock-SI on unsynchronized nodes. The performance turned out to be unacceptable. Therefore, in all the follow-up experiments, we always performed clock synchronization before each test of Clock-SI. It is well known that the accuracy of clock synchronization depends on network latency. In a LAN, the error of clock synchronization can achieve sub-millisecond. In a WAN, the error is around tens of milliseconds. In an initial test, we manually varied the time skew among the nodes, to see how the error affects the performance of Clock-SI. As shown in Figure 6, Clock-SI’s performance drops dramatically when the time skew increases. So does its abort rate. Therefore, in the experiments, we used two versions of Clock-SI – one is completely synchronized, representing the LAN setting (denoted by Clock0); the other represents the WAN setting (denoted by Clock20), to which we introduce a random time skew between -20ms and 20ms.

C. **Performance on Standard Benchmarks**

Our first set of experiments aimed to find out the general performance of the various schedulers on the benchmarks of TPC-C and SmallBank. In the experiments, we varied the number of slaves participating in the test, to study the scalability of the various schedulers. We also varied the proportion of distributed transactions (from 20% to 50%), to see how it influences the overall performance. Figures 7-10 show the performance of the various schedulers.

As we can see, the scalability of conventional SI is the worst among all the schedulers. The growth of its performance clearly slows down when the number of the nodes reaches 16. At this point, the master is about to be saturated by the requests from the slaves, as each transaction requires two
rounds of communication with the master. In contrast, the optimal scheduler performs the best, as it requires the least inter-node communication. PostSI, CV, DSI and Clock-SI all outperform conventional SI, while none of them can beat the optimal scheduler.

SmallBank contains simpler and shorter transactions than TPC-C. In our experiments, it also incurs less contention on data. As a result, the system’s throughput on SmallBank is much higher than that on TPC-C. For schedulers such as DSI and conventional SI, a large proportion of their overheads is caused by the communication with the central coordinator, which occurs more frequently when transactions are shorter and run faster. Therefore, they perform worse in SmallBank than in TPC-C. In contrast, the schedulers of Clock-SI, PostSI and CV are less affected by the length of transactions.

The performance of Clock-SI can approaches that of the optimal scheduler, if the clocks of different nodes are completely synchronized. However, when there is time skew, its performance drops dramatically. Time skew hinders its performance in two ways – first, the latency increases on the nodes that are ahead of the time, as they need to wait for the nodes that fall behind; second, the abort rate increases on the nodes that fall behind, because they often have to read older versions of the data. Especially when there is high contention, time skew can cause very high abort rate. As a result, Clock20 performs the worst in TPC-C. Through an extra set of experiments, we found that Clock-SI’s performance approaches that of PostSI if the time skew is limited to 5ms. In other words, Clock5 should perform as well as PostSI. However, controlling the time skew does not seem trivial. True time devices, such as GPS clocks and atomic clocks, can be used to minimize the time skew and ensure the performance and stability of Clock-SI. This is out of the scope of this paper.

The performance of DSI is mainly influenced by distributed transactions. First, every distributed transaction needs to communicate with the central coordinator, which can be eventually saturated by an increasing number of requests. Second, DSI’s abort rate on distributed transactions is usually high, as a mismatch between a local timestamp and the global timestamp can cause abort. This is confirmed by its authors in [13]. As shown in the results of SmallBank, when there are a large number of short distributed transactions, the scalability of DSI starts to hit a wall. In the experiments of TPC-C, as the transactions are significantly lengthier, the bottleneck of coordination is not yet visible. Therefore, DSI’s performance is similar to that of PostSI on TPC-C.

PostSI and CV outperform DSI and the Clock20 in most of our experiments. The overhead of PostSI is mainly caused by the communication for negotiating the time intervals of contending transactions. The overhead of CV is mainly caused by the transmission of anti-dependency information among the nodes. As the experiment results show, such extra communication is usually limited, as it only occurs when the contention between distributed transactions is high. Both PostSI and CV appear to scale well in TPC-C and SmallBank. CV performs slightly better than PostSI, as it is weaker than PostSI in isolation and allows for more concurrency.

For transaction management on an MPP platform, the frequency of cross-node communication seems to be one of the dominant factors for performance. The abort rates can also provide an insight about the cause of bad performance. Figure 11 shows the communication cost and the abort rates of the various schedulers in a TPC-C test. It is consistent the performance results in Figure 7.

\[ \text{D. Characteristics of the SI Schedulers} \]

We designed several additional sets of experiments to study the characteristics of the various SI schedulers. We used Smallbank as the benchmark. We varied the degree of contention, the lengths of transactions and the fraction of distributed transactions in Smallbank, and observed how these factors influence the performance of the schedulers. All the experiments were conducted on 20 nodes.

In the first set of experiments, we varied the degree of contention by varying the proportion of transactions that access hotspot data. (On each node, 20 out of the 1 million data items are classified as the hotspot.) Figure 12 shows how the various SI schedulers behave. As expected, when contention increases, the throughput of all the schedulers declines.
which incurs extra cross-node communication. Faster than the others. This is because CV needs to lookup the
schedulers drops. CV’s performance seems to drop slightly communication will occur. Thus, the performance of all the
distributed transactions, more cross-node dependency among transactions. Through the concept of
visibility, we proved the soundness of our approaches. We implemented the PostSI and CV schedulers. Our experimental
evaluation demonstrated their suitability for shared-nothing architectures. Our future research will explore the methods to
enable replication, so that our schedulers can be utilized by a full-fledged distributed database.

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