Exploit Lock Violation for Fault-tolerant Distributed Database System

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***Abstract*—Modern distributed database systems scale horizontally by sharding its data on a large number of nodes. To achieve fault tolerance, most of such systems build their transactional layers on a replication layer, which employs a consensus protocol to ensure data consistency. Synchronization among replicated state machines thus becomes a major overhead of transaction processing. Without a careful design, synchronization could amplify transactions' lock duration and impair the system's scalability. Speculative techniques, such as Controlled Lock Violation (CLV) and Early Lock Release (ELR), prove to be effective in shortening lock duration and boosting performance of transaction processing. An intuitive idea is to use these techniques to optimize geo-replicated distributed databases. In this paper, we show that direct application of speculation is often unhelpful in a distributed environment. Instead, we introduce Distributed Lock Violation (DLV), a speculative technique for geo-replicated distributed databases. DLV minimize the cost of to conduct lock violation, so that it can achieve good performance without incurring severe side effects.**

***Index Terms*—Database System, Distributed Transaction, Locking, High Availability**

I. Introduction

Modern distributed database system scale out by partitioning data across multiple nodes, so that it can run transactions on multiple servers in parallel to increase throughput. When a transaction needs to access multiple partitions, it has to employ a coordination protocol to ensure atomicity. It is commonly known that such distributed transactions can lead to significant performance degradation. This is mainly due to the following reasons [1]:

1. Coordinated commit requires a chatty protocol (e.g., Two-Phase Commit), which introduces tremendous network overheads;

2. Message transmission overlaps with the critical path of transaction commit, which worsens the contention among transactions.

These issues can be more serious for geo-replicated databases, which face increased network overheads. The replication layer of a geo-replicated database often uses a Paxos-like consensus protocol to ensure consistency among replicas. This introduces a significant amount of message transmission, which further increase the duration of transactions. To facilitate replication, we usually split data into small chunks and replicate each chunk independently. As a result, cross-chunk transactions (rather than cross-partition transactions) all become distributed transactions. This makes distributed transaction even more inevitable.

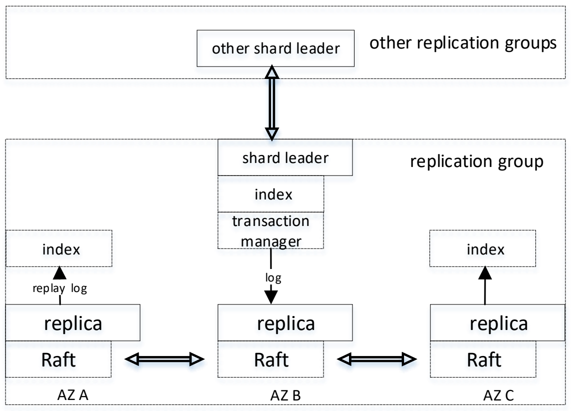


Fig. 1: Typical architecture of geo-replicated distributed DBMS

Fig 1 presents the typical architecture of geo-replicated distributed databases. The database partitions its data into a number chunks. For each chunk, the replication layer replicates its data to several Availability Zones (AZ) [2]. Among the available zones, the replication layer employ a consensus protocol to shield data consistency.

As previous work has discussed [1] [3] [4], this architecture has to rely on a chatty protocol (which integrates the commit protocol and the consensus protocol) to ensure the correctness of transactions. It may fail to scale when confronted with highly contended workload. Nevertheless, this architecture supports a wide range of transaction processing methods. Most industrial data systems choose this two-layer architecture, including Google Spanner [5] [6], NuoDB [7], CockroachDB [8], TiDB [9], etc.

Looking more closely, the main issue is that the commit and consensus protocols enlarge the timespan of the critical paths in transaction processing. This significantly amplifies the overhead of contention. Fig 2 shows the message flow of a

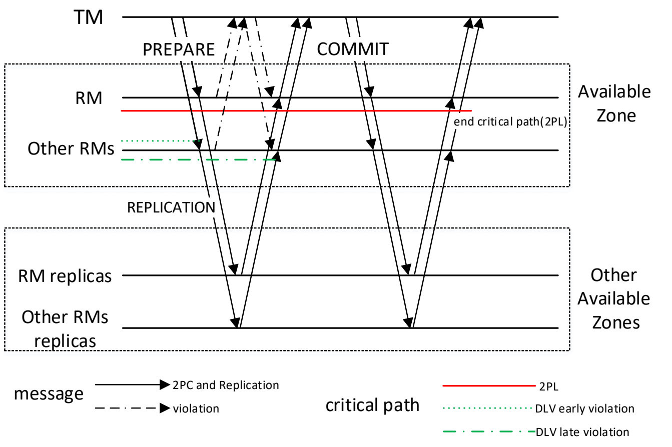


Fig. 2: The message flow and lock holding critical path when the DBMS uses S2PL for concurrency control and 2PC as the commit protocol. The dash arrow lines represents the messages introduced by Distributed Lock Violation (DLV). The red lines represent the critical paths of S2PL. The green dash lines represent the critical paths of DLV.

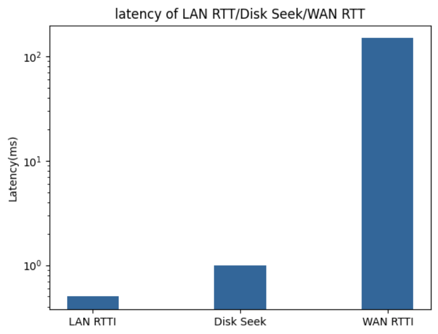


Fig. 3: approximate latency comparison [12]

distributed transaction in such an architecture. We assume that it uses S2PL for concurrency control and 2PC as the commit protocol, and the replication layer is deployed over a WAN. When a transaction requests to commit, the *TM (transaction manager)* issues a 'prepare' message to each *RM (resource manager)*. Then, each *RM* replicates its vote ('prepare commit' or 'prepare abort') to the corresponding replicas through a consensus protocol, before it sends its vote to *TM*. After the *TM* collects all the votes of the *RMs*, it broadcasts the final decision ('commit' or 'abort') to all the *RMs* [1] . Once a *RM* receives the final decision from the *TM*, it replicates the decision to the corresponding replicas. After the consensus of commit or abort is reached, the *RM* can release the locks that it had retained over the accessed data.

We depict lock duration by red lines in Fig 2. As we can see, the lock duration span multiple message round trips, including those over the WAN. This will severely impair the concurrency of transaction processing in face of a high degree of contention.

Speculative techniques, such as Early Lock Release (ELR) [10] and Controlled Lock Violation (CLV) [11], prove to be effective in optimizing centralized transaction processing. These techniques can be extended to a distributed environment.

Speculative techniques improve concurrency by excluding logging from lock duration, Transactions on geo-replicated distributed databases require much more time to ship and persist logs. Fig 3 shows the WAN RTT , LAN RTT and disk seek approximate time costs comparison. This provides more opportunities for exploiting speculation. However, application of speculation in a distributed environment is complex. A number of design choices need to be considered. For instance, when working with 2PC, we need to decide to violate (or release) lock at which phase. As previous work showed [11], lock violation at the first phase may enable a higher degree of concurrency. However, dependency tracing and cascade abort may incur excessive overheads. On the contrary, lock violation at the second phase costs less in dependency tracing and cascade abort, while it has to sacrifice a certain amount of concurrency. On the other hand, not all transactions can benefit from CLV or ELR, e.g., workload with little conflict. It unnecessary and even harmful to apply speculation to all cases.

[1]Depending on variants of implementations, the TM can choose to persist its decision on its log or not.

In this paper, we propose a technique called Distributed Lock Violation (DLV) to enhance the performance of geo-replicated distributed DBMS. DLV decides the best time to perform lock violation by looking at runtime statistic information. It introduces much less penalty compared to previous implementation [11].

The remainder of this paper is organized as follows. Section II reviews related work. Section III shows that a strict schedule is not necessary and hurt the performance of distributed DBMS. Section IV introduces DLV. Section V evaluates DLV and compare it against previous work. Section VI concludes this paper.

II. Relate Work

A. Transaction Processing on Replicated Databases

Most replicated databases rely on state-machine replication(SMR) to realize high availability. SMR often uses a consensus protocol to synchronize different replicas. Synchronization introduces significant amount of network traffic and becomes a major overhead of replicated database systems. Paxos [13] [14] is the most well known consensus protocol. To reach a consensus on a single data update in Paxos, it costs two message round trips, one for choosing a proposal and another for proposing the entry. Multidecree Paxos [15] elects a leader as the only proposer to eliminate the first message round trip. Raft [16] is a similar consensus protocol to Paxos. As it is more understandable, it is widely used in modern database systems. In Raft, it costs at least one message round trip to reach a consensus. Google Spanner [5] [6], NuoDB [7], CockroachDB [8] and TiDB [9] are all geo-replicated DBMSes built upon Paxos or Raft. They all face heavy costs incurred by data synchronization.

To minimize the cost of synchronization, a number of techniques have also been proposed. VoltDB [17] and Calvin [1] employ a deterministic transaction model to reduce the coordination cost of distributed transactions. Deterministic scheduling enables active replication, which allows transactions to kickoff synchronization at the earliest possible time. Tapir [3] relaxes the consistency requirements of the replication layer, so that it can reduce the message round trips to reach consensus. Janus [4] aims to minimize wide-area message round trips by consolidating the concurrency control mechanism and the consensus protocol. It uses deterministic serializable graph tracing to ensure atomicity of transactions. In a nutshell, Tapir and Janus both co-designed the transaction and replication layers of distributed databases, so that they only need to incur one wide-area message round trip to commit a transaction. VoltDB, Calvin, Tapir and Janus all impose strict constraints on the implementation of the transaction layer, making them incompatible with existing systems, such as Spanner and CockroachDB. In contrast, this work focuses on the optimization opportunities in the general implementation of geo-replicated databases (as depicted in Fig 1).

B. *Optimization on Locking based Concurrency Control*

Two-phase locking (2PL) is the most widely used concurrency control mechanism. As a pessimistic method, 2PL assumes a high likelihood of transaction confliction. It uses locks to enforce the order of conflicting transactions. Strict 2PL (S2PL) is a brute force implementation of 2PL. It requires a transaction to preserve all its lock until it ends. As S2PL can easily guarantee transactions' recoverability, many databases choose to use it. When extending S2PL to distributed databases, the lock holding time will be substantially enlarged, as the commit critical path will involve a number of message round trips.

All S2PL implementations adopt a certain approach to resolve deadlocks. In the *no-wait* [18] approach, a transaction immediately aborts if it fails to place a lock. Previous works showed that this is a scalable approach in a distributed environment [19] [18], as it eliminates blocking completely. However, it works poorly when dealing with workload of high contention. Another approach is *wait-die* [20]. It avoids some false-positive aborts encountered by *no-wait* by utilizing timestamps. The *Deadlock detection* approach [21] detects deadlock by explicitly tracing wait-for relationship among transactions. Many centralized database systems [22] [23] adopt this approach, as can deal with contention better. However, deadlock detection in a distributed environment is highly costly, making it the least favorable approach in our case.

To optimize the performance of locking based concurrency control, speculation can be used. Early lock release (ELR) [24] [25] [26] [10] [27] is a typical speculative technique. ELR allows a transaction to release its locks before its commit log is flushed to disk. It was first proposed by DeWitt et al. [24]. Soisalon-Soininen et al. [25] analyzed its correctness in various settings. Many other works applied and evaluated ELR in different types of systems [26] [10] [10] [26]. However, previous works on ELR were limited to centralized database systems.

Control lock violation (CLV) [11] is a more general speculative technique than ELR. It allows certain transactions to ignore certain locks, instead of releasing a lock completely. CLV has been tested on distributed databases. The results show that it can optimize both phases of two-phase commit. CLV needs to trace dependency among transactions. In [11], the authors use a Register and Report (RARA) approach [28] to implement the tracer. RARA work well on a centralized database. In a distributed environment, dependency tracing becomes much more costly. Cascade abort is another side effect of speculation. It can also leads to severe performance downgrade for distributed databases.

III. Scheduling with Lock Violation

In the following, we describe the basic idea of scheduling with lock violation.

A. *Preliminaries and Assumptions*

Our scenario is a geo-replicated distributed database, which shards its data by primary keys. Each data chunk is replicated across a number of remotely located AZs. Physiological logs are transmitted among the AZs to keep the data synchronized. The physiological logs record the row-level write operations of each transaction. A consensus protocol to used to prevent the synchronization from going wrong. We assume that there is a replica leader for each data chunk, which is responsible for coordinating the synchronization process.

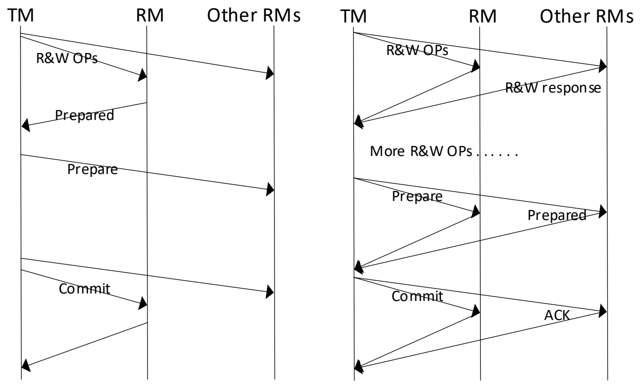
We assume that two types of distributed transactions can be conducted on the system. They are known as one-shot and interactive transactions. Fig 4 shows the message flow during the commit phase of the two types of transactions. We can see that an interactive transaction costs more message round trips than a one-shot transaction, as the TM has to explicitly notify the RMs that they can prepare to commit.

B. *Concepts of Scheduling*

Before presenting the idea of lock violation, we first review the conceptual model of transaction processing. Most of the contents can be found in previous work, such as [20].

1) Transaction and History: Suppose that a distributed database resides on n nodes, represented by $R = \{r_1, r_2, ...r_n\}$. A transaction T_i can run on any m of the n nodes ($1 \leq m \leq n$), represented by $S = \{s_1, s_2, ...s_m\} \subseteq R$, The transaction T_i is composed of a series of operations, where each can be a read or a write or other command operation including abort, commit, etc. Let $r_i[x]$ denote that T_i reads the record x. Let $w_i[x]$ denote that T_i writes on the record x. Let c_i, a_i, p_i^c, p_i^a denote the commit, abort, prepare-to-commit and prepare-to-abort operations of T_i respectively.

A transaction history is a collection $H = \{h_1, h_2, ..., h_n\}$, in which each h_u is the local history of H on the node s_u, which is a sequence of operations issued by different transactions.



(a) one-shot transaction, which does not need to explicitly ask the RMs to prepare

(b) interactive transaction, which has to notify each RM to prepare

Fig. 4: Commit message flows for one-shot and interactive distributed transactions

2) Deterministic and Non-deterministic Abort: A number of reasons can cause a transaction to abort. In general, they can be categorized as:

1. User requested abort. These are aborts specified by the program logic (e.g., requiring a transaction to abort if it has accessed non-existent records).

2. Violation of isolation (e.g., serializability). These are aborts commanded by the database system.

3. Database node failure. For simplicity, we assume fail-stop failures only.

We call the first two types of abort *deterministic abort* and the last one *non-deterministic abort*. In a database system, deterministic aborts occur much more frequently than non-deterministic abort. Therefore, the prices we are willing to pay for the two kinds of abort are very different.

3) Dependencies among Transaction: There are there kinds of data dependency among transactions, known as *wr-dependency*, *ww-dependency* and *rw-dependency*. In a local history h, if T_j reads T_i's write on x, we call it a *write-after-read(wr) dependency* and denote it by $w_i[x] \rightarrow r_j[x]$. Analogically, if T_j overwrites T_i's write on x, it is called a *write-after-write(ww) dependency* and denoted by $w_i[x] \rightarrow w_j[x]$. If T_i reads x before T_j writes on x, it is called a *read-after-write(rw) dependency* and denoted by $r_i[x] \rightarrow w_j[x]$.

Based on data dependencies, we can define *commit dependency*. T_j has a *commit dependency* on T_i, written as $T_i \rightarrow T_j$, if T_j cannot commit prior to T_i. In other words, T_i aborts, T_j has to abort too. T_j has a *commit dependency* on T_i, if and only if T_j has a *rw-dependency* on T_i.

4) Relaxation on Strictness: Most traditional database systems adopt Strict Two-Phase Locking (S2PL) [29] to ensure the recoverability of transactions. Strictness implies that a transaction cannot read a previous write by another transaction which has not committed yet.

Strictness is not a necessary condition for a correct schedule. Although it simplifies the implementation of a database system, it sacrifices the concurrency of transaction processing. For instance, the schedule H_1 of Fig 5 is serializable and strict. (The data dependencies include $r_1[x] \rightarrow w_3[x]$, $w_1[x] \rightarrow r_2[y]$, $r_2[y] \rightarrow w_3[x]$, which are not cyclical.) In contrast, H_2 in Fig 6 is serializable but not strict. In Fig 6, there are three records x, y, z, residing on s_1, s_2, s_3 respectively. T_1 writes y and x. T_2 reads T_1's write on x before T_1 commits. Transaction T_3 overwrites T_2's write before T_2 commits. As we can see, by violating strictness, H_2 enables a higher degree of concurrency.

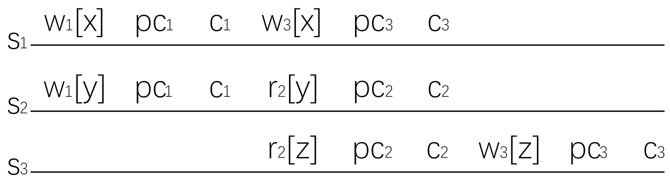

S1 w1[x] pc1 c1 w3[x] pc3 c3
S2 w1[y] pc1 c1 r2[y] pc2 c2
S3 r2[z] pc2 c2 w3[z] pc3 c3

Fig. 5: A strict and serializabile schedule H_1

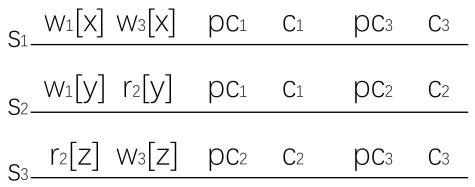

S1 w1[x] w3[x] pc1 c1 pc3 c3
S2 w1[y] r2[y] pc1 c1 pc2 c2
S3 r2[z] w3[z] pc2 c2 pc3 c3

Fig. 6: A non-strict but serializabile schedule H_2

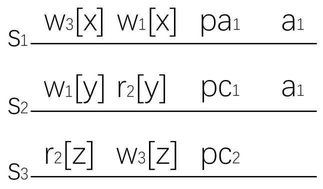


Fig. 7: Schedule H_3, T_1 abort due to non-serializabile

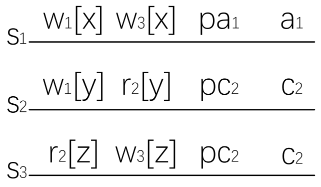


Fig. 8: Schedule H_4, T_2 commit ahead T_1, non-recoverable anomaly

The drawback of strictness is amplified in a geo-replicated and distributed database. The commit of a transaction in such a database usually involve a lot of work, including multiple rounds of communication across the WAN. As strictness requires a transaction to hold locks until it finishes committing, this means a substantially lengthened lock holding time. This will hurt the performance of distributed transactions severely. Our basic idea is to develop a serializable but non-strict scheduler to shorten the lock holding time.

Relaxation on strictness will, however, complicate the recovery mechanism. In a traditional recovery mechanism, such as ARIES, the dependencies among transactions completely comply with the persisting order of their logs. This guarantees the recoverability of scheduling, which requires that a transaction cannot commit before the transactions it has commit dependencies on. When strictness no longer holds, we need other tactics to enforce the commit order, which may incur additional overheads.

C. Lock Violation and Dependency Tracing

Lock violation is a general technique to enable a non-strict scheduler. Its basic idea is to allow a transaction to speculatively access data locked by other transactions. If such speculative data accesses cause a problem, the system should be able to rectify it.

For lock violation to be correct, it should first preserve serializability. Considering the non-strict schedule H_3 in Fig 7, it contains three data dependencies:

$$w_3[x] \rightarrow_s w_1[x],\ w_1[y] \rightarrow_s r_2[y]\ r_2[z] \rightarrow_s w_3[z]$$

This leads to a circle $T_1 \rightarrow T_2 \rightarrow T_3 \rightarrow T_1$, which makes the schedule non-serializable. H_3 is not possible if we apply S2PL. It becomes possible when lock violation is permitted – we allow T_2 to read from uncommitted T_1, and T_1 to overwrite uncommitted T_3. Therefore, to preserve serializability, we need to prevent the data dependencies from forming a dependency cycle. Many techniques can be adopted to achieve this, including the *wait-die* strategy mentioned in the related work.

Secondly, a schedule generated by lock violation must be recoverable. In the schedule H_4 in Fig 8, due to lock violation, T_2 reads T_1's write and commits before T_1. In this case, if T_1 later on decides to abort, it is no longer possible to reverse T_2. Therefore, H_4 is not a recoverable schedule. To ensure recoverability, we need to trace the commit dependencies (i.e., wr-dependencies) among transactions, and force the commit order to comply with the commit dependencies. For instance, if we know that T_2 depends on T_1, we can hold T_2's commit request until T_1 finishes. If T_1 decides to abort, we abort T_2 too. This is known as cascade abort.

In a geo-replicated distributed database, lock violation can potentially be allowed at two occasions in a transaction. The first occasion is when the work of the transaction has been completed successfully and the transaction is ready to commit. We call lock violation at this occasion *late lock violation*. It is only intended to reduce the contention in the commit phases of transactions. The second occasion is before the transaction is ready to commit. We call lock violation at this occasion *early lock violation*. The early lock violation can also fall into two categories: that which violation when an RM begins to decides its prepare decision, and violation when a transaction is even on-going in this RM.

While early lock violation allows us to harness more concurrency, it may cause additional issues. First, early lock violation can result in non-serializable schedules. This has been illustrated by H_3 in Fig 7. In contrast, *late lock violation* is safe. If a transaction can only violate the locks of transactions that are ready to commit, it cannot result in cyclic data dependencies. Second, early lock violation faces a much higher chance of cascade abort. In late lock violation, only non-deterministic aborts can cause cascade abort, as no lock can be violated before deterministic aborts (since the transaction has already been permitted to commit). In early lock violation, all aborts can cause cascade abort. As both types of lock violation have their advantages, our mechanism of Dynamic Lock Violation (DLV) integrates them to achieve the best results.

Secondly, a schedule generated by lock violation must be recoverable. In the schedule H_4 in Fig 8, due to lock violation, T_2 reads T_1's write and commits before T_1. If T_1 decides to abort, it is no longer possible to reverse T_2. As a result, schedule H_4 is not recoverable. To ensure recoverability, we need to trace the wr-dependencies among transactions, to ensure that the commit order complies with the wr-dependencies.

Whether late or early lock violation is adopted, we need to trace the commit dependencies among transactions and make sure they commit in a right order. In a geo-replicated distributed database, the scheduler should follow the following rules to enforce the commit order. Given that there is a commit dependency from T_j to T_i,

1) T_j's RM can persist its 'prepare' state (including replication of the state) only if T_i has committed;
2) T_j's TM can send out 'commit' request only if T_i has committed;
3) If T_i aborts, T_j must also abort too.

IV. Implementation of DLV

In this section, we present how to implement Distributed Lock Violation (DLV). The following contents would include: How DLV can avoid complex recovery algorithm and maintain the most limited amount of dependencies; How DLV choose the proper time of enabling violation; The wait-die policy of DLV use; The pseudocode code description finally.

A. No UNDO Log and Speculative Versions

As mentioned earlier, a non-strict scheduler may require a more complex recovery mechanism for a DB use undo log. For instance, consider the following non-strict schedule:

$$H_5 = w_1[x]w_2[x]a_1a_2$$

If transaction T_1 aborts, it forces T_2 to abort too. This is known as cascade abort. To conduct abort, traditional database systems use undo logs to cancel out a transaction's write operations. Undo becomes a bit tricky when non-strictness is allowed. For instance, let $exp(H_5)$ be an extended schedule of H_5 with undo operations, in which $w_i^-[x]$ denotes that T_i undoes its write on x.

$$exp(H_5) = w_1[x]w_2[x]w_1^-[x]c_1w_2^-[x]c_2$$

To understand why this schedule is wrong and how it can occur, lets take Table I as an example. Suppose the initial value of records x and v of are both 0. The value of records x and its undo log formatted after executing every operations in

$exp(H)$ is shown in Table I. Firstly, T_1 write x, and create its undo log by x's previous value. Later, transaction T_2 violates T_1's lock on x after T_1 write its value 1. T_2 overwrite T_1's write on x and construct its undo log by T_1's write value. Transaction T_1 abort due to some reason and undo its write. Then T_2 also abort and undo its write by a wrong value of x, which is T_2 read from T_1. Finally, after the execution of this schedule, both transaction T_1 and T_2 aborts. The value of x is 1, which the correct result should be the initial value 0.

operations	x	undo
$w_1[x=1]$	1	x=0
$w_2[x=2]$	2	x=1
$w_1^-[x=0]$	0	
c_1	0	
$w_2^-[x=1]$	1	
c_2	1	

TABLE I: x, y values, undo log after the execution of $exp(H_5)$

In other words, when performing cascade abort, we cannot abort each transaction independently. Instead, we have to follow the reserve order of the write operations to perform the recovery. For schedule H_5, a correct recovery expandation may be:

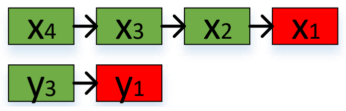
$$exp(H_5) = w_1[x]w_2[x]w_2^-[x]c_2w_1^-[x]c_1$$

This problem is intrinsic with non-strictness. Even if we choose to perform *late lock violation* only, which does not permit lock violation before deterministic abort, cascade abort can still be triggered by non-deterministic abort. To tackle this problem, we need a more complex recovery algorithm, such as SOT [30], than traditional ARIES [31]. The schdeuler must rollback transaction by the reserve order of write operation. To reduce the complexity, DLV maintains uncommitted data versions only in memory and applies the no-steal policy to data persistence. The no-steal policy forbids the system from writing uncommitted data to permanent storages. As a drawback, it may consume more memory space. However, this is not a big concern, considering that today's database servers are usually equipped with large RAM, . Based on this design, we do not need undo log anymore. If a transaction aborts, we only need to discard its uncommitted data in memory.

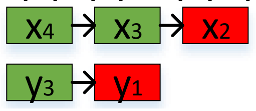
The no-steal policy has another advantage in geo-replicated databases. That is, it pushes data synchronization among replicas to the commit phase, so that through lock violation we can exclude data synchronization completely from the lock holding time.

The version list is structured from the newest version to the oldest version and the last version of this list is the committed version. Since speculation versions are uncommitted data, they are always stored in main memory and needs no persistence. And in permanent storage, there is only single version, the committed version. When a transaction commit, it cooperatively clean its outdated version from the version list. If a transaction would abort, it only needs to remove its write versions from the speculative version list.

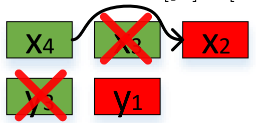
Previously, we have discussed that a *ww* dependency does not affect recoverability. *Late lock violation*, since it has promised serializability, so it can ignore ww and rw dependencies and only trace wr dependencies for recoverability. Fig 9 show a series of schedule access on two contention rows, x, y. The green rectangles are speculative versions and the red ones are committed versions. Although there is ww dependency $w_6[x] \rightarrow w_4[x]$. The abort of T_4 does not cause T_6 cascade abort.



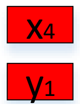
(a) $H_a = w_1[x_0]w_1[y_0]c_1w_2[x_1]w_3[x_2]w_3[y_1]w_4[x_3]$



(b) $H_b = H_ac_2r_5[y_3]w_6[x_4]$



(c) $H_c = H_ba_3a_5$



(d) $H_d = H_cc_4r_6[y_1]c_6$

Fig. 9: speculative(green) and committed(red) versions, x_i express this version is from transaction transaction T_i's write. So, $w_j[x_i]$ express transaction T_j overwrite T_i's write.

B. Lock Violation Occasion

In a distributed transaction, TM assigns work to RM. Once the work is done, TM applies 2PC to end the transaction. In 2PC, both TM and RM need to synchronize with their replicas, to make sure that node failures will not corrupt the data. Synchronization usually takes a long time, especially when the replicas are located remotely. DLV aims to prevent data synchronization from coinciding with the lock holding time.

In *late lock violation*, a transaction allows its lock to be violated only after all RMs confirm that they can commit. Therefore, it requires an additional message round trip between the TM and the RMs before the commit starts. In *early lock violation*, each RM can autonomously decide to when to permit lock violation. It saves a message round trip, but faces a higher chance of cascade abort. We implement 4 types of DLV strategies which are different on viollation time. The lock violation time in *DLV0*, *DLV1*, *DLV1*, *DLV2* is following:

DLV0: once a transaction has finished an access operation, the lock which protects this operation can be violated;

DLV1: when a transaction in an RM has decides its prepare status, the locks on this shard hold by this transaction can be violated;

DLV1x: a transaction ensures all the RM related has decides their prepare, all the transactions's lock can be violated;

DLV2: a transaction enter 2 phase of 2PC, all of the transactions's lock can be violated.

DLV0 and *DLV1* is early lock violation, and *DLV1x* and *DLV2* is late lock violation. DLVx implementation *late lock violation* by adding a message round trip to prevent deterministic abort when use lock violation. This additional message flow shows in Fig 2, the message is show as dotted arrow lines. Before an RM decides to replicate its prepare log, it also sends a *Ready* message to TM and tells TM it will prepare this transaction. *Ready* message shows that the RM will prepare commit or prepare abort. When the *TM* collect all *RM*s *Ready* message, it sends *Violate* messages to tells every *RMs* make their locks are violative. An interactive-transaction can combine these messages with the last operations and prepare requests in passing, as Fig 4b shows. So, the interactive transaction needs no dynamic decides the lock violation time. Interactive transactions use DLV would only work as *late violation*. We mainly focus on discuss one-shot transaction model here. For a one-shot transaction, this message flow also takes less time than the overall message flow of 2PC because of no log replication time dealy. Especially when all the RMs and the TM are located in LAN, there is no WAN message RTT.

C. Deadlock Prevention and Dependency Tracing

DLV adopts the *wait-die* protocol to prevent deadlocks and dependency cycles. A transaction is assigned a unique timestamp when it starts. Conflicting transactions can then be ordered by their timestamps. Once a transaction attempts to access a data item locked by a transaction with a larger timestamp, the transaction is aborted. As mentioned earlier, this is a simple and effective approach in a distributed environment. The conflict transaction operations are queued base their transaction id's order. In early lock violation mode, DLV would consolidating all the wait in the final wait. Ealy lock violation also obeys *wait-die* rule. A transaction can only violate a lock with greater transaction id, which is the same as wait rule. By this approach, transaction dependencies graph can not create a circle. It can reduce possible false positive abort when exploit early lock violation.

DLV applies the register-and-report method [28] to trace commit dependencies among transactions. Each transaction maintains an in-dependency count (in attribute of Context in Algorithm 1), which records how many transactions it depends on. At the same time, each transaction also maintains an out-dependency transaction set (out attribute of Context in Algorithm 1), which records all the transactions depending on it. When a transaction T reads from S, DLV registers the commit dependency by adding T to the out set of S and incrementing in of T by one. A transaction cannot start committing if it's in value is greater than 0, which means some of its in-dependency transaction has not committed yet. When a transaction aborts, it may cause its out-dependency transactions abort cascadely. For this purpose, it notifies all its out-dependency transactions to abort too, by setting their in values to a negative value, for example $-\infty$. This means that the transaction must abort. When a transaction commits, this transaction will traverse its out set and decrease its out-dependency transactions' in values by one. This dependency information can be maintained in main memory. If there is node failure, this information can be dropped safely.

D. Pseudocode Description

Algorithm 1 Transaction Context and Lock Structure Definination

```
1: struct Context {
2:     xid;              ▷ transaction id
3:     in;               ▷ in dependency transaction count
4:     out;              ▷ out dependency transaction set
5:     locks;            ▷ lock acquired
6:     write_set;        ▷ write row set
7:     violate;          ▷ EARLY or LATE
8: };
1: struct Lock {
2:     xid;              ▷ transaction id
3:     key;              ▷ locked key
4:     mode;             ▷ lock mode, READ or WRITE
5:     violate;          ▷ EARLY, LATE or NONE
6: };
```

Algorithm 1 describes the transaction context on RM and the lock structure. The meanings of the elements are shown in the comments. Both these structures has a violate attribute to indicate which lock violation strategy the transaction used. The *violate* value of Context are initialized when transaction context is created. It would be assigned by EARLY if DLV0 or DLV1 setting, and LATE if DLV1x or DLV2 setting. Lock structure has *violate* attribute which is initilized as NONE to show it cannot be violated. Depends on lock violation strategy, it would becomes EARLY or LATE to express the lock can be violated early or late.

In Algorithm 2, Function Lock decreases the procedure of transaction locking. It has 4 input arguments: transaction context T, the key to lock K, lock mode M and violation type V. First the Lock function check if there's a conflict lock in the lock table. It also test weather it can be violated by the current transaction. Transaction T can only wait a lock who hold by another transaciton has a greater transaction xid[2]. Otherwise, transacion T would fail to acquire this lock and abort later. Once any locks are violated, T would register its dependency if it is necessary. Finally, Lock return the newly created lock.

Algorithm 3 inclues test violable function and enable violation function. Function Violable return weather a transaction T can violate lock L on K in M mode. EnableTxV enable all the locks on a shard can be violated. EnableLkV is used to make a lock can be violated.

[2]we suppose there is no lock upgrade, so there are no conflict locks hold by one transaction

Algorithm 4 depicts the execution phase of a transaction includes Read and Write functions. Read and Write function would lock the read or write key. The lock would be add to the lock set of transaction context if the lock acquire succeed. After the transaction access the key, EnableLkV would be called to enable this lock violable if current database is in DLV0 setting.

Algorithm 5 describes the prepare phase of a transaction. When use DLV1 setting, transaction The transacion context would wait on its in dependency. The transaction context cannot enter prepare phase if its in value is greater than 0, which stands it has in dependency transaction. Once in become 0, the transaction processing can go and continue enter its prepare phase. In some situations, there may be cascade abort. in value less than 0 indicates T would cascade abort.

Algorithm 6 the final phase of 2PC. When a RM receive *Commit* message from TM, it would run Commit function. In Commit function, transaction would test if DLV2 setting on and enable violation. The it clean its write version , write log and release the locks. If current transaction has out dependencies, the in value of the dependency transaction would be decreased one. The transaction reports the dependency transactions stop waiting if in equal 0 after the subtraction. Abort procedure first write the abort log, clean dirty version of this transaction and release log. Transactions who depends on this transacion would be noticed to abort by seting their in dependency count negative.

V. Experiments and Evaluations

We develop a replicated distributed DBMS demo and evaluate implement 4 types of DLV, as we disscuss previously. As a comparison with DLV, we also implement S2PL wait die(S2PL) scheme.

A. Experiments Setting

Our experiments performed on a cluster of 13 Aliyun ecs.g6.large server. In which one node is used to issue transaction requests. Each server has 2 virtual CPU with 2.5GHz clock speed, 8GB RAM, runs Ubuntu 18.04. The data is partitioned by 4 shards, each shard has 3 replicas which is replicated across 3 AZs, which is located at Heyuan, Hangzhou and Zhangjiakou. Every AZ has a full data copy of each shard. The internal network bandwidth of each AZ is 1Gbps. We choose a modifies version TPCC and YCSB workload. All the transactions are distributed transactions. The TPCC data is sharded by the warehouse id. The Item table is replicated to all shards. Each transaction will retry after 3 seconds if it aborts for violation serializability. The evaluation both tested on both scattered (leader) mode and gathered (leader) mode. In gathered mode, all of the replica leaders are located in the same AZes. In scattered mode, the replica leaders are not located in the same AZes.

B. TPCC Performance Evaluation

TPC-C specification models a realword workload. TPC-C models a company has many warehouses and district, customers order their products.

Algorithm 2 Transaction Lock/UnLock a Key
T transaction context
K lock key
L the lock to be violated
M lock mode(*Read* or *Write*)

```
1: function LOCK($T, K, M$)
2:     wait if $\exists$ such $l \in lock\_table$,
       $Conflict(K, T, l, M)$ and
       not $Violable(K, T, l, M)$
3:         if $T.xid > l.xid$
4:             return nil                    ▷ lock die error
5:         end if
6:     end wait
7:     for all such $l \in lock\_table$,
8:     $Violable(K, T, l, M)$ do
9:         $s \leftarrow$ transaction context retrieved by $l.xid$
10:        if $T.violate = EARLY$ or ( $T.violate = LATE$
           and $l.lmode = WRITE$ and $M = READ$ ) then
11:            if $T.xid \notin s.out$ then
12:                $s.out\_set \leftarrow s.out \cup T.xid$
13:                $T.in \leftarrow T.in + 1$
14:            end if
15:        end if
16:    end for
17:    create lock $l' \leftarrow \{T.xid, K, M, None\}$
18:    return lock $l'$
19: end function
1: function UNLOCK($T, K$)
2:     if $\exists l \in lock\_table$, $l.xid = T.xid$ and
3: $l.key = K$ then
4:         remove such $l$ from $lock\_table$
5:     end if
6: end function
1: function CONFLICT($K, T, L, M$)
2:     if $T.xid \neq L.xid$ and $K = L.key$ then
3:         return L.lmode = WRITE or M = WRITE
4:     else
5:         return false
6:     end if
7: end function
```

In our version TPC-C benchmark, We exclude think time of TPCC and only NewOrder procedure is used. The data is partitioned on shards by warehouse id. The item table is replicated by all the shard. We also extends TPC-C workload make all the transactions are distributed transaction by enforce every order has item from different warehouse located different node.

First we simulate log cost on different an LAN environment by adding log write latency. Fig 10 shows the throughput and abort rate of TPCC NewOrder procedure on 4 shards. Every shards replicated transactional log with user setting latency time. The result shows that late violation can perform better when there is a large log replicaiton latency.

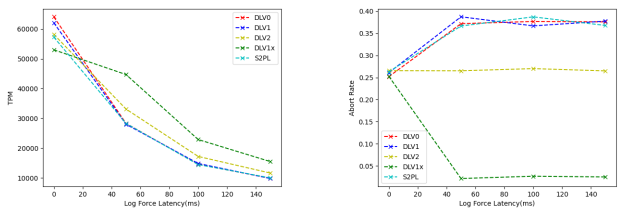
Algorithm 3 Lock Violation Test and Enable Violation

T transaction context
K lock key
L the lock to be violated
M lock mode
VT violation type, EARLY or LATE

```
function VIOLABLE($T, K, L, M$)
    if $CONFLICT(K, T, L, M)$ and
$L.violate \neq NONE$ then
        if $L.violate = EARLY$
and $T.violate = EARLY$ then
            return $T.xid < l.xid$
        else if $L.violate = LATE$ then
            return true
        end if
    else
        return false
    end if
end function
```

```
function ENABLETXV($T, VT$)
    for all $l \in T.locks$ do
        $l.violate \leftarrow VT$
    end for
end function
```

```
function ENABLELKV($L, VT$)
    $l.violate \leftarrow VT$
end function
```



(a) NewOrder TPM (b) NewOrder abort

Fig. 10: throughput and abort rate of different terminal different log write costs

In the following experiments, we all run test on a real replication environment.

Fig 11 is the performance comparation of different warehouse number. DLV perform better when warehouse number is small. The less warehouse number the data divided, the more contention the transaction would incur. DLV only optimize contention wrokload. DLV1x perform better among other approach various on may warehouse number.

Fig 12 shows the NewOrder performance of when adding terminal numbers of each node in the gathered mode. In gathered mode, message between replica leaders has the minimize RTT. DLV1x can benefit from less RTT cost. DLV1x scale well when adding terminal number of each shard. Fig 13 shows that even in scattered mode, leader communication has more cost.

Algorithm 4 Execution Phase of Transaction

T transaction context
K read/write key
V value write

```
function READ($T, K$)
    $l \leftarrow Lock(T, K, READ)$
    if $l$ is $nil$ then
        return lock failed                ▷ $T$ will abort later
    else
        $T.locks \leftarrow T.locks \cup l$
        if DLV0 setting, call $EnableLkV(l, EARLY)$
        return $v$
    end if
end function
```

```
function WRITE($T, K, V$)
    $l \leftarrow Lock(T, K, WRITE)$
    if $l$ is $nil$ then
        return lock failed                ▷ $T$ will abort later
    else
        $T.locks \leftarrow T.locks \cup l$
        assign $K$'s newest version by value $V$
        add tuple to $T.write\_set$
        if DLV0 setting, call $EnableLkV(l, EARLY)$
        return write success
    end if
end function
```

Algorithm 5 Prepare Phase of Transaction T

```
function PREPARE($T$)
    if DLV1 setting, call $EnableTxV(T, EARLY)$
    if DLV1x setting, issue $TM$ a $\{Ready\}$ message.
    wait if $T.in > 0$
    if $T.in < 0$ then
        write prepare abort log
        response TM message $\{Prepare\ Abort\}$ ▷ cascade
    else if $T.in = 0$ then
        write prepare commit log
        response TM message $\{Prepare\ Commit\}$
    end if
end function
```

DLV1x perform better than other DLV strategies.

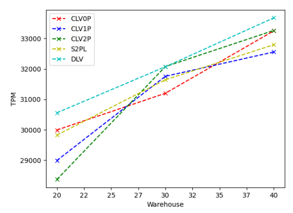
C. YCSB Performance Evaluation

In YCSB workload,

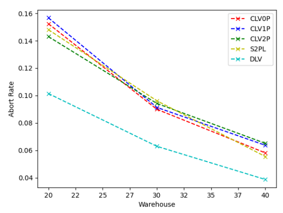
We evaluate different data skew.

VI. Conclution

We extend CLV to distributed transaction and evaluate its performance on a geo-replicated environment. Our distributed version CLV, i.e. DLV, can dynamically decide to violate lock at the most suitable time. DLV merge many discrete waits at transaction running time into one final wait when commit. According to our evaluation, DLV can improve performance of contention workload for shortening critical path. DLV

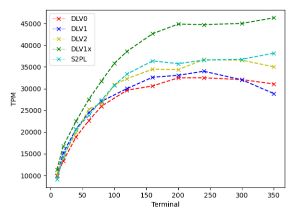


(a) NewOrder TPM

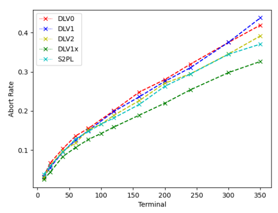


(b) NewOrder abort

Fig. 11: throughput and abort rate of different warehouse numbers

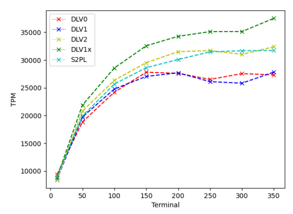


(a) NewOrder throughput

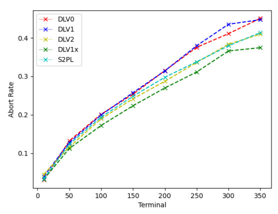


(b) NewOrder abort rate

Fig. 12: throughput and abort rate when adding terminal number of each shard, in gathered mode

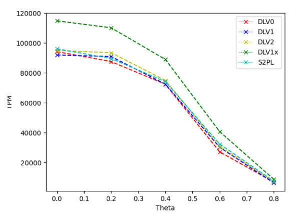


(a) NewOrder throughput

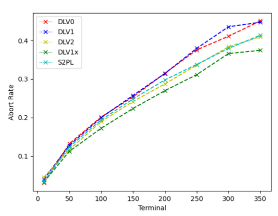


(b) NewOrder abort rate

Fig. 13: throughput and abort rate when adding terminal number of each shard, in scattered mode



(a) YCSB work throughput



(b) YCSB workload abort rate

Fig. 14: throughput and abort rate when adding terminal number of each shard, in scattered mode

Algorithm 6 Commit/Abort Phase of Transaction T

```
1:  function COMMIT($T$)
2:      if CLV2 setting, call $EnableTV(T)$
3:      write commit log
4:      for all $r \in T.write\_set$ do
5:          make the version written by $T$ becomes commit version of tuple $t$
6:      end for
7:      for all $l \in T.locks$ do
8:          UNLOCK($l$)
9:      end for
10:     for $t \in T.out$ do
11:         $t.in \leftarrow t.in - 1$                ▷ keep exactly once
12:         if $t.in = 0$ then
13:             report $t.in = 0$                    ▷ stop waiting on $t$
14:         end if
15:     end for
16:     response TM message {Commit ACK}
17: end function
1:  function ABORT($T$)
2:      write abort log
3:      for all $r \in T.write\_set$ do
4:          remove the version written by $T$ in $r$
5:      end for
6:      for all $l \in T.locks$ do
7:          UNLOCK($l$)
8:      end for
9:      for $t \in T.out$ do
10:         $t.in \leftarrow -\infty$
11:         report $t.in < 0$                        ▷ stop waiting on $t$
12:     end for
13:     response TM message {Abort ACK}
14: end function
```

can adapt to different workloads. It minimize unnecessary dependency tracing cost and cascade abort penalty against previous work.

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