Scalable Replication in Database Clusters

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Abstract. In this paper, we explore data replication protocols that provide both fault tolerance and good performance without compromising consistency. We do this by combining transactional concurrency control with group communication primitives. In our approach, transactions are executed at only one site so that not all nodes incur in the overhead of producing results. To further reduce latency, we use an optimistic multicast technique that overlaps transaction execution with total order message delivery. The protocols we present in the paper provide correct executions while minimizing overhead and providing higher scalability.

1 Introduction

Conventional algorithms for database replication emphasize consistency and fault tolerance instead of performance [BHG87]. As a result, database designers ignore these algorithms and use *lazy* replication instead, thereby compromising both fault-tolerance and consistency [GHOS96]. A way out of this dilemma [KA98,KA] is to combine database replication techniques with group communication primitives [HT93]. This approach has produced efficient *eager* replication protocols that guarantee consistency and increase fault tolerance. However, in spite of some suggested optimizations [KPAS99,PS98], this new type of protocols still have two major drawbacks. One is the amount of redundant work performed at all sites. The other is the high abort rates created when consistency is enforced. In this paper, we address these two issues. First, we present a protocol that minimizes the amount of redundant work. Transactions, even those over replicated data, are executed at only one site. The other sites only install the final changes. With this, and unlike in existing replication protocols, the aggregated computing power actually increases as more nodes are added. This is a significant advantage in environments with expensive transaction processing (e.g., dynamic web pages). A negative aspect of this protocol is that it might abort transactions in order to guarantee serializability. To reduce the rate of aborted transactions while still providing consistency, we propose a second protocol based on a transaction reordering technique.

The paper is organized as follows. Section 2 introduces the system model. Sections 3 and 4 describe the algorithms. Section 5 discusses fault tolerance aspects. Section 6 contains the correctness proofs. Section 7 concludes the paper.

2 System Model

In a replicated database, a group of nodes $N = \{N_1, N_2, ..., N_n\}$, each containing the entire database, communicate by exchanging messages. Sites only fail by crashing (no byzantine failures) and there is always at least one node available.

2.1 Communication Model

The system uses various group communication primitives [HT93]. Regarding message ordering, we use a multicast primitive not providing any order, a primitive providing *FIFO order* (messages of one sender are delivered in FIFO order) and one providing a *total order* (all messages are delivered at all sites in the same order). In regard to fault-tolerance, we use both a *reliable delivery service* (whenever a message is delivered at an available site it will be delivered at all available sites) and a *uniform reliable delivery service* (whenever a message is delivered at any faulty or available site it will be delivered at all available sites). We assume a virtual synchronous system, where all group members perceive membership (view) changes at the same virtual time, i.e., two sites deliver exactly the same messages before installing a new view.

We use an aggressive version [KPAS99] of the optimistic total order broadcast presented in [PS98]. Each message corresponds to a transaction. Messages are optimistically delivered as soon as they are received and before the definitive ordering is established. With this, the execution of a transaction can overlap with the calculation of the total order. If the initial order is the same as the definitive order, the transactions can simply be committed. If the final order is different, additional actions have to be taken to guarantee consistency. This optimistic broadcast is defined by three primitives [KPAS99]. *To-broadcast*(m) broadcasts the message m to all the sites in the system. *Opt-deliver*(m) delivers message m optimistically to the application (with no order guarantees). *To-deliver*(m) delivers m definitively to the application (in a total order). This means, messages can be opt-delivered in a different order at each site, but are to-delivered in the same total order at all sites. A sequence of opt-delivered messages is a *tentative order*. A sequence of to-delivered messages is the *definitive order* or total order. Furthermore, this optimistic multicast primitive ensures that every to-broadcast message is eventually opt-delivered and to-delivered by every site in the system. It also ensures that no site to-delivers a message before opt-delivering it.

2.2 Transaction Model

Clients interact with the database by issuing transactions, i.e., partially ordered sets of read and write operations. Two transactions conflict if they access the same data item and at least one of them is a write. A history H of committed transactions is serial if it totally orders all transactions. Two histories H_1 and H_2 are conflict equivalent, if they are over the same set of transactions and order conflicting operations in the same way. A history H is serializable, if it is conflict equivalent to some serial history [BHG87]. For replicated databases,

the correctness criterion is 1-copy-serializability [BHG87]. Using this criterion, each copy must appear as a single logical copy and the execution of concurrent transactions must be equivalent to a serial execution over all the physical copies.

In this paper, concurrency control is based on *conflict classes* [KPAS99]. Each conflict class represents a partition of the data. Transactions accessing the same conflict class have a high probability of conflicts, as they can access the same data, while transactions in different partitions do not conflict and can be executed concurrently. In [KPAS99] each transaction must access a single *basic* conflict class (e.g., C_x). We generalize this model and allow transactions to access *compound conflict classes*. A compound conflict class is a non-empty set of basic conflict classes (e.g., $\{C_x, C_y\}$). We assume that the (compound) conflict class of a transaction is known in advance. Each site has a queue CQ_x associated to each basic conflict class C_x. When a transaction is delivered to a site, it is added to the queues of the basic conflict classes it accesses. This concurrency control mechanism is a simplified version of a lock table [GR93].

Each conflict class has a *master* site. We use a *read-one/write-all available* approach. Queries (read only transactions) can be executed at any site using a snapshot of the data (i.e., they do not interfere with update transactions). Update transactions are broadcast to all sites, however they are only executed at the master site of their conflict class. We say a transaction is *local* to the master site of its conflict class and is *remote* everywhere else.

3 Increasing Scalability

3.1 The Problem and a Solution

The scalability of data replication protocols heavily depends on the update ratio. To see why, consider a centralized system capable of processing t transactions per second. Now assume a system with n nodes, all of them identical to the centralized one. Assume that the fraction of updates is w. Assume the load of local transactions at a node is x transactions per second. Since nodes must also process the updates that come from other nodes, the following must hold: $x + w\,(n-1)\,x = t$, that is, a node processes x local transactions per second, plus the percentage of updates arriving at other nodes ($w\,x$) times the number of nodes. From here, the number of transactions that can be processed at each node is $x = t\,(1 + w\,(n-1))^{-1}$. The total capacity of the system is n times that expression which yields, with t normalized to 1, $n\,(1 + w\,(n-1))^{-1}$. This expression has a maximum of n when $w = 0$ (there are no updates) and a minimum of 1 when $w = 1$ (all operations are updates).

Thus, as the *update factor* w approaches 1, the total capacity of the system tends to that of a single node, independently of how many nodes are in the system. Note that the drop in system capacity is very sharp. For 50 nodes, $w = 0.2$ (20% updates) results in a system with a tenth of the nominal capacity.

This limitation can be avoided if transactions execute only at one site (the local site) and the other sites only install the corresponding updates. This requires significantly less than actually running the transactions as has been shown

in [KA00]. In order to guarantee consistency, the total order established by the to-delivery primitive is used as a guideline to serialize transactions. All sites see the same total order for update transactions. Thus, to guarantee correctness, it suffices for a site to ensure that conflicting transactions are ordered according to the definitive order. Transactions can be executed in different orders at different sites if they are not serialized with respect to each other.

When an update transaction T is submitted, it is multicast to all nodes. This message contains the entire transaction and it is first opt-delivered at all sites which can then proceed to add the corresponding entries in the local queues. Only the local site executes T: whenever T is at the head of any of its queues the corresponding operation is executed on a shadow copy of the data. With this, aborting a transaction simply requires to discard the shadow copies. When the transaction commits the shadow copies become the valid versions of the data.

When a transaction is to-delivered at a site, the site checks whether the definitive and tentative orders agree. If they agree, the transaction can be committed after its execution has completed. If they do not agree, there are several cases to consider. The first one is when the lack of agreement is with non-conflicting transactions. In that case, the ordering mismatch can be ignored. If the mismatch is with conflicting transactions, there are two possible scenarios. If no local transactions are involved, the transaction can simply be rescheduled in the queues before the transactions that are only opt-delivered but not yet to-delivered. With this, to-delivered transactions will then follow the definitive order. If local transactions are involved, the procedure is similar but local transactions (that have been executed in the wrong order) must be aborted and rescheduled again (by putting them back in the queues in the proper order).

Once a transaction is to-delivered and completely executed the local site broadcasts the commit message containing all updates (also called write set WS). Upon receiving a commit message (which does not need any ordering guarantee), a remote site installs the updates for a certain basic conflict class as soon as the transaction reaches the head of the corresponding queue. When all updates are installed the transaction commits.

3.2 Example

Assume there are two basic conflict classes C_x, C_y and two sites N and N'. N is the master of conflict classes $\{C_x\}$, and $\{C_x, C_y\}$. N' is the master of $\{C_y\}$. We denote the conflict class of a transaction T_i by C_{T_i}. Assume there are three transactions, $C_{T_1} = \{C_x, C_y\}$, $C_{T_2} = \{C_y\}$ and $C_{T_3} = \{C_x\}$. That is, T_1 and T_3 are local at N and T_2 is local at N'. The tentative order at N is: T_1, T_2, T_3 and at N' is: T_2, T_3, T_1. The definitive order at both sites is: T_1, T_2, T_3. When all the transactions have been opt-delivered, the queues at each site are as follows:

At N:	At N':
$CQ_x = T_1, T_3$	$CQ_x = T_3, T_1$
$CQ_y = T_1, T_2$	$CQ_y = T_2, T_1$

At site N, T_1 can start executing both its operations on C_x and C_y since it is at the head of the corresponding queues. When T_1 is to-delivered the orders

are compared. In this case, the definitive order is the same as the tentative order and hence, T_1 can commit. When T_1 has finished its execution, N will send a commit message with all the corresponding updates. N can then commit T_1 and remove it from the queues. The same will be done for T_3 even if, in principle, T_2 goes first in the final total order. However, since these two transactions do not conflict, this mismatch can be ignored. Parallel to this, when N receives the commit message for T_2 from N', the corresponding changes can be installed since T_2 is at the head of the queue CQ_y. Once the changes are installed, T_2 is committed and removed from CQ_y.

At site N', T_2 can start executing since it is local and at the head of its queue. However, when T_1 is to-delivered, N' realizes that it has executed T_2 out of order and will abort T_2, moving it back in the queue. T_1 is moved to the head of both queues. Since T_3 is remote at N', moving T_1 to the head of the queue CQ_x does not require to abort T_3. T_1 is now the first transaction in all the queues, but it is a remote transaction. Therefore, no transaction is executing at N'. When the commit message of T_1 arrives at N', T_1's updates are applied, T_1 commits and is removed from both queues. Then, T_2 will start executing again. When T_2 is to-delivered and completely executed, a commit message with its updates will be sent, and T_2 will be removed from CQ_y.

3.3 The NODO Algorithm

The first algorithm we propose, NODO (NOn-Disjoint conflict classes and Optimistic multicast), follows that in [KPAS99]. The algorithm is described according to the different phases in a transaction's execution: a transaction is opt-delivered, to-delivered, completes execution, and commits. We assume access to the queues is regulated by locks and latches [GR93]. There are some restrictions on when certain events may happen. For instance, a transaction can only commit when it has been executed and to-delivered. Waiting for the to-delivery is necessary to avoid conflicting serialization orders at the different sites. Each transaction has two state variables to ensure this behavior: The *execution state* of a transaction can be *active* (as soon as it is queued) or *executed* (when its execution has finished). A transaction can only become executed at its master site. The *delivery state* can be *pending* (it has not been to-delivered yet) or *committable* (it has been to-delivered). When a transaction is opt-delivered its state is set to active and pending. In the following we assume that whenever a transaction is local and the first one in any of its queues, the corresponding operations are submitted for execution.

We assume that each of the phases is done in an atomic step. For instance, adding a transaction to the different queues during opt-delivery or rescheduling transactions during to-delivery is not interleaved with any other action. Note that aborting a transaction simply involves discarding the shadow copy, the transaction itself is kept in the queues but in different positions.

Upon **Opt-delivery** of T_i
  Mark T_i as active and pending
  For each conflict class $C_x \in C_{T_i}$
    Append T_i to the queue CQ_x
  EndFor

Upon **TO-delivery** of T_i:
  Mark T_i as committable
  If T_i is executed **then**
    Broadcast commit (WS_{T_i})
  Else *(still active or not local)*
    For each $C_x \in C_{T_i}$
      If First(CQ_x) = T_j
        $\wedge$ Local(T_j)
          $\wedge$ Pending(T_j) **then**
        Abort T_j
        Mark T_j as active
      EndIf
      Schedule T_i before the first
      pending transaction in CQ_x
    EndFor
  EndIf

Upon **complete execution** of T_i
  If T_i is marked as committable **then**
    Broadcast commit(WS_{T_i})
  Else
    Mark T_i as executed
  EndIf

Upon receiving **commit(WS_{T_i})**
  If $\neg$ Local (T_i) **then**
    Delay until T_i becomes committable
    For each $C_x \in C_{T_i}$
      When T_i becomes the first in CQ_x
        Apply the updates of WS_{T_i}
          corresponding to C_x
        Remove T_i from CQ_x
      EndFor
    Else
      Remove T_i from all C_{T_i}
    EndIf
    Commit T_i

4 Reducing Transaction Aborts

In the NODO algorithm, a mismatch between the local optimistic order and the total order may result in a transaction being aborted. The resulting abort rate is not necessarily very high since for this to happen, the transactions must conflict, appear in the system at about the same time, and the site where the mismatch occurs must be the local site where the aborted transaction was executing. In all other cases there are no transaction aborts, only reschedulings. Nevertheless, network congestion and high loads can lead to messages not being spontaneously ordered and, thus, to higher abort rates. The number of aborted transactions can be reduced by taking advantage of the fact that NODO is a form of master copy algorithm (remote sites only install updates in the proper order). Thus, a local site can unilaterally decide to change the serialization order of two local transactions (i.e., follow the tentative order instead of the definitive total order), thereby avoiding the abort. To guarantee correctness, the local site must inform the rest of the sites about the new execution order (by appending this information to the commit message). Special care must be taken with transactions that belong to a non basic conflict class (e.g., $C_{T_i} = \{C_x, C_y\}$). A site can only follow the tentative order $T_1 \rightarrow_{OPT} T_2$ instead of the definitive order $T_2 \rightarrow_{TO} T_1$, if T_1's conflict class C_{T_1} is a subset of T_2's conflict class C_{T_2} and botho are local transactions. Otherwise, inconsistencies could occur. We call this new algorithm REORDERING as the serialization order imposed by the definitive order might be changed for the tentative one.

4.1 Example

Assume a database with two basic conflict classes C_x and C_y. Site N is the master of the conflict classes $\{C_x\}$ and $\{C_x, C_y\}$. N' is the master of conflict class $\{C_y\}$. To show how reordering takes place, assume there are three transactions $C_{T_1} = C_{T_3} = \{C_x, C_y\}$, and $C_{T_2} = \{C_x\}$. All three transactions are local to N. The tentative order at both sites is T_2, T_3, T_1. The definitive order is T_1, T_2, T_3. After opt-delivering all transactions they are ordered as follows at both sites:

QC_x : T_2, T_3, T_1
QC_y : T_3, T_1

At site N, T_2 and T_3 can start execution (they are local and are at the head of one of their queues). Assume that T_1 is to-delivered at this stage. In the NODO algorithm, T_1 would be put at the head of both queues which can only be done by aborting T_2 and T_3. This abort is, however, unnecessary since N controls the execution of these transactions and the other sites are simply waiting to be told what to do. Thus, N can simply decide not to follow the total order but serialize according to the tentative order. This is possible because all transactions involved are local and the conflict classes of T_2 and T_3 are a subset of T_1's conflict class. When such a reordering occurs, T_1 becomes the *serializer transaction* of T_2 and T_3. T_2 does now not need to wait to be to-delivered to commit. Being at the head of the queue and with its serializer transaction to-delivered, the commit message for T_2 can be sent once T_2 is completely executed (thereby reducing the latency for T_2). The commit message of T_2 also contains the identifier of the serializer transaction T_1. The same applies to T_3.

Site N' has at the beginning no information about the reordering. Thus, not knowing better, when T_1 is to-delivered at N', N' will reschedule T_1 before T_2 and T_3 as described in the NODO algorithm. However, when N' receives the commit message of T_2, it realizes that a reordering took place (since the commit message contains the information that T_2 has been serialized before T_1). N' will then reorder T_2 ahead of T_1 and mark it committable. N', however, only reschedules T_2 when T_1 has been to-delivered in order to ensure 1-copy serializability. The rescheduling of T_3 will take place when the commit message for T_3 arrives, which will also contain T_1 as the serializer transaction. In order to prevent that T_2 and T_3 are executed in the wrong order at N', commit messages are sent in FIFO order (note, that FIFO is not needed in the NODO algorithm).

As this example suggests, there are restrictions to when reordering can take place. To see this, consider three transactions $C_{T_1} = \{C_x\}$, $C_{T_2} = \{C_y\}$ and $C_{T_3} = \{C_x, C_y\}$. T_1 and T_3 are local to N, T_2 is local to N'. Now assume that the tentative order at N is T_3, T_1, T_2 and at N' it is T_1, T_2, T_3. The definitive total order is T_1, T_2, T_3. After all three transactions have been opt-delivered the queues at both sites look as follows:

Queues at site N:	Queues at site N':
QC_x : T_3, T_1	QC_x : T_1, T_3
QC_y : T_3, T_2	QC_y : T_2, T_3

Since T_3 is local and it is at the head of its queues, N starts executing T_3. For the same reasons, N' starts executing T_2. When T_1 is to-delivered at N, T_3 cannot

be reordered before T_1. Assume this would be done. T_3 would commit and the commit message would be sent to N'. Now assume the following scenario at N'. Before N' receives the commit message for T_3 both T_1 and T_2 are to-delivered. Since T_2 is local, it can commit when it is executed (and the commit is sent to N). Hence, by the time the commit message for T_3 arrives, N' will produce the serialization order $T_2 \rightarrow T_3$. At N, however, when it receives T_2's commit, it has already committed T_3. Thus, N has the serialization order $T_3 \rightarrow T_2$, which contradicts the serialization order at N'.

This situation arises because $C_{T_1} = \{C_x, C_y\}$ is not a subset of $C_{T_3} = \{C_x\}$ and, therefore, T_1 cannot be a serializer transaction for T_3. In order to clarify why subsets (i.e., the conflict class of the reordered transaction is a subset of the conflict class of the serializer transaction) are needed for reordering, assume that T_1 also accesses C_y (with this, $C_{T_3} \subseteq C_{T_1}$). In this case, the queues are:

Queues at site N:	Queues at site N':
QC_x : T_3, T_1	QC_x : T_1, T_3
QC_y : T_3, T_1, T_2	QC_y : T_1, T_2, T_3

The subset property guarantees that T_1 conflicts with any transaction with which T_3 conflicts. Hence, T_1 and T_2 conflict and N' will delay the execution and commitment of T_2 until the commit message of T_1 is delivered. As the commit message of the reordered transaction T_3 will arrive before the one of T_1, T_3 will be committed before T_1 and thus before T_2 solving the previous problem. This means, that both N and N' will produce the same serialization order $T_3 \rightarrow T_1 \rightarrow T_2$.

4.2 REORDERING Algorithm

In general, the REORDERING algorithm is similar to NODO except in a few points (in the following we omit the actions upon opt-delivery since they are they same as in the NODO algorithm). The commit message must now contain the identifier of the *serializer transaction* and follow a FIFO order. As in NODO, when a transaction T_i is to-delivered, the transaction is marked as committable. At T_i's local site, any non to-delivered local transaction T_j whose conflict class C_{T_j} is a subset of C_{T_i} and that precedes T_i in the queues (reorder set RS) is marked as committable (since now the commit order is no longer the definitive but the tentative order). Thus, it is possible that when a reordered transaction is to-delivered the transaction is already marked as committable or even has been committed. In this case the to-delivery message is ignored. Local non to-delivered conflicting transactions that cannot be reordered and have started execution are aborted (abort set, AS). When the to-delivered transaction is remote, the algorithm behaves as the NODO algorithm. Note that a remote reordered transaction T_i cannot commit at a site until its serializer transaction is to-delivered at that site. When this happens, T_i is rescheduled before its serializer transaction. The rescheduling together with the FIFO ordering ensure that remote transactions will commit at all sites in the same order in which they did at the local site.

Upon **to-delivery** of transaction T_i
If $\neg$ Committed(T_i) $\wedge$ Pending(T_i) **then**
(T_i has not been reordered)
If Local(T_i) **then**
If T_i is marked executed **then**
Broadcast commit (WS_{T_i,T_i})
Else *(T_i has not finished yet)*
$AS = \{T_j | C_{T_j} \cap C_{T_i} \neq \emptyset \wedge C_{T_j} \not\subseteq C_{T_i}$
$\wedge \exists C_x \in C_{T_j} \cap C_{T_i}$: T_j = First(CQ_x)
$\wedge$ Pending(T_j) $\wedge$ Local(T_j)$\}$
For each $T_j \in AS$
(abort conflicting transactions that cannot be reordered)
Abort T_j and mark it as active
EndFor
(try to reorder transactions)
$RS = \{T_j | C_{T_j} \subseteq C_{T_i} \wedge T_j \rightarrow_{opt} T_i$
$\wedge$ Pending(T_j) $\wedge$ Local(T_j)$\}$
For each $T_j \in RS \cup \{T_i\}$
in opt-delivery order
Mark T_j as committable
Mark T_i as serializer (Ser) for T_j
Schedule T_j before the first pending transaction in all $CQ_x | T_j \in C_x$
EndFor
EndIf
Else *(It is a remote transaction)*
Mark T_i committable
For each conflict class $C_x \in C_{T_i}$
If T_j = First(CQ_x) $\wedge$ Pending(T_j)
$\wedge$ Local(T_j) **then**
Abort T_j and mark it as active
EndIf
Schedule T_i before the first transaction marked as pending in queue CQ_x
EndFor
EndIf
Else *(transaction has been reordered)*
Ignore the message
EndIf

Upon **complete execution** of T_i
If T_i is marked as committable **then**
Broadcast commit $WS_{T_i,Ser(T_i)}$
Else
Mark T_i as executed
EndIf

Upon receiving **commit** WS_{T_i,T_j}
If not Local(T_i) **then**
Delay until T_j is committable
If $T_i \neq T_j$ **then**
Mark T_i as committable
EndIf
EndIf
For each $C_x \in C_{T_i}$
If not Local(T_i) **then**
If $T_i \neq T_j$ **then**
Reschedule T_i just before T_j in CQ_x
EndIf
When T_i becomes the first in CQ_x
apply the updates of WS_{T_i,T_j} corresponding to C_x
EndIf
Remove T_i from CQ_x
EndFor
Commit T_i

5 Dealing with Failures

In our system, each site acts as a primary for the conflict classes it owns and as a backup for all other conflict classes. In the event of site failures, the available sites simply have to select a new master for the conflict classes of the failed node. The new master will also take over the responsibility for all pending transactions

for which the failed node was the owner (i.e., where the commit message has not been received by the available sites). Such a master replacement algorithm guarantees the availability of transactions in the presence of failures. That is, a transaction will commit as far as there is at least one available site.

For both algorithms, transaction messages must be uniformly multicast because only then it is guaranteed that the master will only execute and commit a transaction when all sites will receive it, and thus, be able to take over if the master crashes (reliable multicast does not provide this since the master can commit a transaction which the other sites have not yet received).

In the NODO algorithm, commit messages do not need to be uniform. Local transactions can even be committed before multicasting the commit message. The worst that can happen is that a master commits a transaction and fails before the commit message reaches the other sites. When a new master takes over, it will reexecute the transaction and send a new commit message. As the total order is always followed inconsistencies cannot arise.

In the REORDERING algorithm, commit messages must be uniform and the master may not commit the transaction before the commit message is delivered. If the commit message were not uniform, a master could reorder a transaction, send the commit message and then crash. If the rest of the replicas do not see the commit message, they would use a different serialization order (as the failed node's optimistic order is unknown to the other sites).

6 Correctness

In this section we prove the correctness (i.e., 1-copy-serializability), liveness, and consistency of the protocols. The proofs assume histories encompassing several group views. Important for both protocols is the fact that transactions are enqueued (respectively rescheduled) in one atomic step. Hence, there is no interleaving between transactions and all sites produce automatically serializable histories. As a result, in order to prove 1-copy-serializability, it suffices to show that all histories are conflict equivalent. Since conflict equivalence requires histories to have the same set of transactions, we refer in the corresponding proofs only to the available sites.

6.1 Correctness of NODO

We will show that all sites order conflicting transactions according to the definitive total order.

Definition 1 (Direct conflict). *Two transactions T_1 and T_2 are in direct conflict if they are serialized with respect to each other, $T_1 \longrightarrow T_2$, and there are no transactions serialized between them: $\nexists T_3 \mid T_1 \longrightarrow T_3 \longrightarrow T_2$.*

Lemma 1 (Total order and Serializability in NODO). *Let H_N be the history produced at site N, let T_1 and T_2 be two directly conflicting transactions in H_N. If $T_1 \longrightarrow_{TO} T_2$ then $T_1 \longrightarrow_{H_N} T_2$.*

Proof (lemma 1): Assume the lemma does not hold, i.e., there is a pair of transactions T_1, T_2 such that $T_1 \longrightarrow_{H_N} T_2$ but $T_2 \longrightarrow_{TO} T_1$. The fact that T_2 precedes T_1 in the total order means that T_2 was to-delivered before T_1. Since T_1 and T_2 are in direct conflict, there was at least one queue where both transactions had entries. If $T_1 \longrightarrow_{H_N} T_2$, then the entry for T_1 must have been ahead in the queue. However, upon to-delivery of T_2, if T_1 was the first transaction, NODO would have aborted T_1 and rescheduled it after T_2. If T_1 was not the first in the queue, NODO would have put T_2 ahead of T_1 in the queue. In both cases this would result in $T_2 \longrightarrow_{H_N} T_1$ which contradicts the initial assumption. □

Lemma 2 (Conflict equivalence in NODO). *For any two sites N and N', H_N is conflict equivalent to $H_{N'}$.*

Proof: (lemma 2) ¿From Lemma 1, all pairs of directly conflicting transactions in both H_N and $H_{N'}$ are ordered according to the total order. Thus, H_N and $H_{N'}$ are conflict equivalent since they are over the same set of transactions and order conflicting transactions in the same way. □

Theorem 1 (1CPSR in NODO). *The* NODO *algorithm produces 1-copy-serializable histories.*

Proof: (theorem 1) Since the histories of all available nodes are conflict equivalent (lemma 2) and serializable, the global history is 1-copy-serializable. □

6.2 Liveness of NODO

Theorem 2 (Liveness in NODO). *Each to-delivered transaction T_i eventually commits in the absence of catastrophic failures.*

Proof: (theorem 2) The theorem is proved by induction.
Induction Basis: Let T_i be the first to-delivered transaction. Upon to-delivery, each site places T_i at the head of all its queues. Thus, T_i's master can execute and commit T_i, and then multicast the commit message. Remote sites will apply the updates and also commit T_i.
Induction Hypothesis: The theorem holds for the to-delivered transactions with positions $n \leq k$, for some $k \geq 1$, in the definitive total order, i.e., all transactions that have at most $k-1$ preceding transactions will eventually commit.
Induction Step: Assume that transaction T_i is at position $n = k+1$ in the definitive total order when it is to-delivered. Each node places T_i in the corresponding queues after any committable transaction (to-delivered before T_i) and before any pending transaction (not yet to-delivered). All committable transactions that are now ordered before T_i have lower positions in the definitive total order. Hence, they will all commit according to the induction hypothesis and be removed from the queues. With this, T_i will eventually be the first in each of its queues and, as in the induction basis, eventually commit.

In all cases, if the master fails before the other sites have received the commit, the new master will reexecute T_i and resend the commit message. □

6.3 Consistency of NODO

Failed sites obviously do not receive the same transactions as available sites. Let $\mathcal{T}$ be the subset of transactions to-delivered to a node before it failed.

Theorem 3 (Consistency of failed sites with NODO). *All transactions, T_i, $T_i \in \mathcal{T}$, that are committed at a failed node N are committed at all available nodes. Moreover, the committed projection of the history in N is conflict equivalent to the committed projection of the history of any of the available nodes when this history is restricted to the transactions in $\mathcal{T}$.*

Proof: (theorem 3) A transaction T_i can only commit at N when it is to-delivered. Since we use uniform reliable delivery, T_i will also be to-delivered and known at all available sites. If T_i was not local at N, then N must have received a commit message from T_i's master. If this master is available for sufficient time all other available sites will also receive the commit message. If the master fails a new master will take over, execute T_i and resend the commit. This procedure will repeat if the new master also fails before the rest of the system receives the commit message. Since we assume there are some available nodes, eventually one of these nodes will become the master and the transaction will commit. If the transaction was local at N, the same argument applies. The equivalence of histories follows directly from Lemma 2. □

6.4 Correctness of REORDERING

In the REORDERING algorithm it is not possible to use the total order as a guideline since nodes can reorder local transactions. Thus, we start by proving that transactions not involved in a reordering cannot get in between the serializer and the transaction being reordered. Let T_s be the serializer transaction of the transactions in the set $\mathcal{T}_{T_s}$.

Lemma 3 (Reordered). *A reordered transaction T_i is always serialized before its serializer transaction T_s, that is, if $T_i \in \mathcal{T}_{T_s}$ then $T_i \longrightarrow T_s$.*

Proof (lemma 3): It follows trivially from the algorithm. □

Lemma 4 (Serializer in REORDERING). *For all transactions T_i, $T_i \in \mathcal{T}_{T_s}$ there is no transaction T_j, $T_j \notin \mathcal{T}_{T_s}$, such that $T_i \longrightarrow T_j \longrightarrow T_s$.*

Proof (lemma 4): Assume that N is the master site where the reordering takes place. Since T_s is the serializer of T_i, $T_i \longrightarrow_{OPT} T_s$, and $T_s \longrightarrow_{TO} T_i$. Additionally, from Lemma 3 $T_i \longrightarrow T_s$. There are two cases to consider: (a) $T_j \longrightarrow_{TO} T_s$ and (b) $T_s \longrightarrow_{TO} T_j$.
(a): since T_j is to-delivered before T_s, in the queues T_j is before T_i, and T_i is before T_s. With T_j ahead of their queues, T_i and T_s cannot be committed until T_j commits. Thus, T_j cannot be serialized in between T_i and T_s.
(b): since T_s is to-delivered before T_j and $T_j \notin \mathcal{T}_{T_s}$, all sites will put T_s ahead of T_j in the queues (T_j cannot have committed because it has not yet been

to-delivered), if it was not the case. Since $C_{T_i} \subseteq C_{T_s}$, this effectively prevents transactions from getting in between T_i and T_s. Any transaction T_j trying to do so will conflict with T_s and since T_s has been to-delivered before T_j, T_j has to wait until T_s commits. By that time, T_i will have committed at its master site and its commit message will have been delivered and processed at all sites before the one of T_s. Therefore, the final serialization order will be $T_i \longrightarrow T_s \longrightarrow T_j$. □

Lemma 5 (Conflict Equivalence in REORDERING). *For any two sites N and N', H_N is conflict equivalent to $H_{N'}$.*

Proof: (lemma 5) We show that two directly conflicting transactions T_1 and T_2 with conflict classes C_{T_1} and C_{T_2} are ordered in the same way at N and N'. We have to distinguish several cases:

• $C_{T_1} \subseteq C_{T_2}$, T_1 and T_2 have the same master N'', and $T_2 \longrightarrow_{TO} T_1$:

(a) If N'' reorders T_1 and T_2 with respect to the total order, then, from Lemma 4, no transaction $T_i \notin \mathcal{T}_{T_2}$ can be serialized in between. The commit for T_1 will be sent before the commit for T_2 in FIFO order. Hence, all sites will then execute T_1 before T_2.

(b) If N'' follows the total order to commit T_1 and T_2, then other sites cannot change this order. The argument is similar to that in Lemma 1 and revolves about the order in which transactions are committed at all sites.

• $C_{T_1} \subseteq C_{T_2}$, T_1 and T_2 have the same master N'', and $T_1 \longrightarrow_{TO} T_2$:

(c) If $C_{T_1} = C_{T_2}$ then cases (a) and (b) apply exchanging T_1 and T_2.

(d) Otherwise $C_{T_1} \subset C_{T_2}$. In this case, N'' has no choice but to commit T_1 and T_2 in to-delivery order (the rules for reordering do not apply). From here, and using the same type of reasoning as in Lemma 1, it follows that all sites must commit T_1 before T_2.

• either $C_{T_1} \subseteq C_{T_2}$ and T_1 and T_2 do not have the same master, or $C_{T_1} \cap C_{T_2} \neq \emptyset$ and neither $C_{T_1} \nsubseteq C_{T_2}$ nor $C_{T_2} \nsubseteq C_{T_1}$.

(e) If T_1 or T_2 are involved in any type of reordering at their nodes, Lemma 4 guarantees that there will be no interleavings between the transactions involved in the reordering and the other transaction. Thus, one transaction will be committed before the other at all sites and, therefore, all sites will produce the same serialization order.

(f) If T_1 and T_2 are not involved in any reordering, then similar to Lemma 1, both of them will be scheduled in the same (total) order at all sites and then committed.

• $C_{T_1} \cap C_{T_2} = \emptyset$.

(g) If there is no serialization order between T_1 and T_2 then they do not need to be considered for equivalence.

(h) If there is a serialization order between T_1 and T_2, it can only be indirect. Assume that in N: $T_1 \ldots \longrightarrow T_i \longrightarrow T_{i+1} \longrightarrow \ldots T_2$. Between each pair of transactions in that sequence, there is a direct conflict. Thus, for each pair, the above cases apply and all sites order the pair in the same way. From here it follows that T_1 and T_2 are also ordered in the same way at all sites. □

Theorem 4 (1CPSR in REORDERING). *The* REORDERING *algorithm produces 1-copy-serializable histories.*

Proof: (theorem 4) ¿From Lemma 5, all histories are conflict equivalent. Moreover, they are all serializable. Thus, the global history is 1-copy-serializable. □

6.5 Liveness of REORDERING

Theorem 5 (Liveness in REORDERING). *Each to-delivered transaction T_i eventually commits in the absence of catastrophic failures.*

Proof: (theorem 5) The proof is by induction.
Induction Basis: Let T_i be the first to-delivered transaction. Upon to-delivery, each remote site will place T_i at the head of all its queues. At the local node, there might be some reordered transactions before T_i hence, T_i will be their serializer. All these transactions can be executed and committed, so that T_i will eventually be executed and committed. Remote sites will apply the updates of the reordered transactions and T_i in FIFO order and will also commit T_i.
Induction Hypothesis: The theorem holds for the to-delivered transactions with positions $n \leq k$, for some $k \geq 1$, in the definitive total order, i.e., all transactions that have at most $k-1$ preceding transactions will eventually commit.
Induction Step: Assume that transaction T_i is at position $n = k+1$ in the definitive total order when it is to-delivered. There are two cases:

a) *T_i is reordered.* This means there is a serializer transaction T_j with a position $n \leq k$ in the total order and T_i is ordered before T_j. Since T_j, according to the induction hypothesis, commits and T_i is executed and committed before T_j at all sites, the theorem holds.

b) *T_i is not a reordered transaction.* T_i will be rescheduled after any committable transaction and before any pending transaction. There exist two types of committable transactions rescheduled before T_i.

i. *Not reordered transactions:* They have a position $n \leq k$ and will therefore commit and be removed from the queues according to the induction hypothesis.

ii. *Reordered transactions:* Each reordered transaction that is serialized by transaction $T_k \neq T_i$ will commit before T_k and T_k will commit according to the previous point (i). All transactions $T_j \in \mathcal{T}_{T_i}$ (i.e., T_i is the serializer) are ordered directly before T_i in the queues (Lemma 3). Let T_k be the first not reordered transaction before this set of reordered transactions. T_k will eventually commit according to the previous point (i), and therefore also all transactions in $\mathcal{T}_{T_i}$ and T_i itself.

Failures lead to masters reassignment but do not introduce different cases to the above ones. □

6.6 Consistency of REORDERING

Again, let $\mathcal{T}$ be the subset of transactions to-delivered to a node before it failed.

Theorem 6 (Consistency of failed sites with REORDERING). *All transactions, T_i, $T_i \in \mathcal{T}$, that are committed at a failed node N are committed at all available nodes. Moreover, the committed projection of the history in N, is conflict equivalent to the committed projection of the history of any of the available nodes when this history is restricted to the transactions in $\mathcal{T}$.*

Proof: (theorem 6) Since both transaction and commit messages are sent with uniform reliable multicast, all transactions and their commit messages in $\mathcal{T}$ have been to-delivered to all available sites and can therefore commit at all sites. The equivalence of histories, follows directly from Lemma 5. □

7 Conclusions

In this paper, we have proposed two replication protocols for cluster based applications. These protocols solve the scalability problem of existing solutions and minimize the number of aborted transactions. We are currently implementing and experimentally evaluating the protocols and, as part of future work, we will deploy a web farm with a replicated database built upon these protocols. For this purpose we will use TransLib [JPAB00], a group-based TP-monitor.

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