A Compacted Recovery for Certification-Based Database Replication

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Abstract

Certification-based database replication protocols provide a good basis to develop replica recovery when they provide the snapshot isolation level. For such isolation level, no readset needs to be transferred between replicas nor checked in the certification phase. Additionally, such protocols need to maintain a historic list of writesets that is used for certifying the transactions that arrive to the commit phase. Such historic list can be used to transfer the missed state of a recovering replica. We study the performance of the basic recovery approach –to transfer all missed writesets– and a version-based optimization –to transfer the latest version of each missed item, compacting thus the writeset list–, and the results show that such optimization reduces a lot the recovery time.

1 Introduction

Replication has been the regular solution for achieving high availability. But such level of availability requires that faulty replicas were recovered. Database replication is a special kind of highly-available service since in this case replica recovery implies the application of the missed updates, being impossible a complete state transfer since it needs a long time to be completed. Even transferring only the missed updates, there is no easy way to complete such recovery in a short time.

There have been many good works devoted to database replication recovery [1, 3, 10, 13], but almost none of them has provided a rigorous performance study of the proposed approaches. The aim of this paper is to show that replica recovery is not easy when the load of the replicated system is not light, and that some optimizations can partially overcome such problem. To this end, we have tried to select the database replication kind [18] that provides the best support for developing an easy recovery: certification-based replication. In this replication variant a historic list of the applied writesets needs to be maintained in order to certificate transactions (i.e., validate and locally decide in each replica about the success of each terminating transaction). Such a historic writeset list can be stored and used for transferring the missed updates to recovering replicas. Additionally, the resulting replication protocol does not need any voting termination [19] and provides very good performance if the conflicting rate is low [18]. Moreover, for the snapshot isolation level, a certification-based replication protocol is the natural solution, since it does not demand readset transfers. So, such kind of replication protocol provides an ideal basis to research on replica recovery and a basic recovery protocol can be easily developed.

But such a basic recovery protocol does not provide good performance (i.e., a short recovery time). So, some optimizations are needed in order to get acceptable results. To this end, we have combined a version-based approach, similar to those proposed by other research groups (e.g., in some of the recovery variants of [13]) and in some of our previous papers [3, 4] but specifically adapted to a certification-based replication protocol. Such optimization introduces a negligible overhead and shortens the recov-

ery time to only a 7,5% of its original length, with a medium workload (8 clients in a 4-replica system, with a crash interval that generates 1000 missed writesets), as shown in Section 5.

The rest of this paper is structured as follows. Section 2 presents the assumed system model. Section 3 describes the replication protocol taken as the basis for our recovery proposals. Section 4 thoroughly explains the recovery strategies. Section 5 discusses the performance results. Finally, Section 6 presents some related work and Section 7 gives the conclusions.

2 System Model

We assume a partially synchronous distributed system –where clocks are not synchronized but the message transmission time is bounded– composed by N nodes where each one holds a replica of a given database; i.e., the database is fully replicated in all system nodes. These replicas might fail according to the partial-amnesia crash failure model proposed in [6], since all already committed transactions are able to recover but on-going ones are lost when a node crashes. We consider this kind of failures as we want to deal with node recovery after its failure.

Each system node has a local DBMS that is used for locally managing transactions. On top of the DBMS a middleware is deployed in order to provide support for replication. More information about our MADIS middleware can be found in [12, 14]. This middleware also has access to a group communication service (GCS, on the sequel).

A GCS provides a communication and a membership service supporting virtual synchrony [5]. The communication service features a total order multicast for message exchange among nodes through reliable channels. Membership services provide the notion of view (current connected and active nodes with a unique view identifier). Changes in the composition of a view (addition or deletion) are delivered to the recovery protocol. We assume a primary component membership [5]. In a primary component membership, views installed by all nodes are totally ordered (there are no concurrent views), and for every pair of consecutive views there is at least one process that remains operational in both views. The GCS groups messages delivered in views [5]. The uniform reliable multicast facility [9] ensures that if a multicast message is delivered by a node (faulty or not) then it will be delivered to all available nodes in that view. All these characteristics permit us to know which writesets have been applied in the context of an installed view. In this work, we use Spread [17] as our GCS.

We use a replication protocol based on certification [18], which does not require any kind of voting in order to decide how a transaction should be terminated (either committing or aborting).

3 Replication Protocol

We have selected the SIR-SBD protocol (see Figure 1) described in [14] for a case study of our recovery mechanisms, because it is a good sample of a certification-based [18] database replication protocol, providing the snapshot isolation level [2] and thus avoiding the transfer of transaction readsets.

Initialization:
1. lastvalidated_tid := 0
2. lastcommitted_tid := 0
3. ws_list := $\emptyset$
4. tocommit_queue_k := $\emptyset$

I. Upon operation request for T_i from local client
1. If select, update, insert, delete
   a. if first operation of T_i
      - T_i.start := lastcommitted_tid
      - T_i.priority := 0
   b. execute operation at R_k and return to client
2. else /* commit */
   a. T_i.WS := getwriteset(T_{ik}) from local R_k
   b. if T_i.WS = $\emptyset$, then commit and return
   c. T_i.priority := 1
   d. multicast T_i using total order

II. Upon receiving T_i in total order
1. obtain wsmutex
2. if $\exists\ T_j \in$ ws_list : T_i.start $<$ T_j.end $\wedge$ T_i.WS $\cap$ T_j.WS $\neq \emptyset$
   a. release wsmutex
   b. if T_i is local then abort T_i at R_k else discard
3. else
   a. T_i.end := ++lastvalidated_tid
   b. append T_i to ws_list and tocommit_queue_k
   c. release wsmutex

III. T_i := head(tocommit_queue_k)
1. if T_i is remote at R_k
   a. begin T_{ik} at R_k
   b. apply T_i.WS to R_k
   c. $\forall\ T_j$: T_j is local in $R_k \wedge T_j$.WS $\cap$ T_i.WS $\neq \emptyset$ $\wedge$ T_j has not arrived to step II
      (this is analyzed by our conflict detector, concurrently with the previous step III.1.b)
      - abort T_j
2. commit T_{ik} at R_k
3. ++lastcommitted_tid
4. remove T_i from tocommit_queue_k

Figure 1: SIR-SBD algorithm at replica R_k

This protocol uses an atomic multicast [9], i.e., a reliable multicast with total order delivery, and thus it ensures that the writesets being multicast

by each replica at commit time are delivered in all replicas in the same order. It uses two data structures for dealing with writesets: `ws_list`, which stores all the writesets known (i.e., delivered) until now, and `tocommit` queue, which holds those writesets locally certified but not yet applied in the local database replica. Moreover, for each transaction, the attributes `start` and `end` hold something similar to the transaction start and commit timestamps, respectively. Due to the total order multicast and the behaviour of the protocol, the second counter is the same for a system transaction in all the replicas, i.e., all the replicas identify with the same commit timestamp a system transaction –this will be handy when studying the performance graphs.

Note that we have tacitly assumed that the underlying database system is supposed to be able to check for conflicts, and to abort transactions the access patterns of which violate the snapshot isolation level rules.

This protocol is also based in the existence of a block detection mechanism [14]. We have assigned the following priorities to the transactions. All transactions are initialized with a 0 priority level. They get level 1 when they are multicast in their local node or when their writeset is delivered in their remote nodes. This ensures the correctness of this alternative, since our blocking detection mechanism aborts a transaction only if all of these conditions are satisfied. Otherwise, no particular action is taken:

- The transaction to be aborted is local.
- It has not locally requested its commit; i.e., its writeset has not been multicast.
- The transaction that causes its abortion has been generated for applying a remote writeset.

This approach satisfies the correctness criteria of the snapshot isolation level, since the writeset above mentioned is associated to a transaction that has successfully passed its global validation phase. It already has a commit timestamp which of course is in the range of the [start, commit] interval of the local transaction, since the latter has not yet requested its commit.

4 Recovery Strategies

We describe a basic recovery in Section 4.1 and its optimized version in Section 4.2. The optimization consists in compacting the list of missed writesets, maintaining only the last version of each missed item.

4.1 Basic Recovery

As a general overview of the main goal of our recovery protocol, let us say that one node (recoverer) will transfer the missed writesets to the recovering node arranged by their respective versions. This means that user application transactions executed on the recovering node will run under GSI in a slower replica. As it may be seen there are no restrictions to execute user transactions in the replica and transactions executing at other replicas will behave as it nothing happens in the system. To achieve this we take the ideas outlined in [7].

A recovering replica R_i joins the group, triggering a view change. As part of this procedure, the recovering protocol instance running in R_i multicasts an *ask-for-help* message indicating the $version_i$ of its last applied writeset –this version corresponds to the commit timestamp of the last transaction applied in that node. No message activity in the recovering node is done –all messages delivered are ignored– until this message is delivered. At this moment, the recovering node starts to enqueue the total order delivered messages –with writeset information about other transactions in the system sent by the rest of the replicas– to be processed later.

In parallel to this, a deterministic procedure takes place to choose a recoverer replica. The recoverer replica (R_j), after receiving the *ask-for-help* message, starts a recovery thread that sends a point-to-point message with all the missed writesets starting from $version_i + 1$, i.e., the recoverer node sends the portion of its `ws_list` that covers from $version_i + 1$ to the end of the `ws_list` at that moment. Note that this approach guarantees that the recovering node will receive, on one hand, the writesets from $version_i + 1$ to the last known version of the recoverer at the moment of the *ask-for-help* message reception. On the other hand, the writesets delivered in total order in the system after the *ask-for-help* message will be enqueued in the recovering message buffer. This way, it is ensured that the recovering node will not miss any writeset.

When this point-to-point recovery message is delivered to the recovery protocol, it stores this information in both the `ws_list` and the `tocommit` queue, as all these writesets were already certified in the recoverer node. Then, the replication protocol is ready to directly apply in the database the writesets in the `tocommit` queue and to start certifying its own enqueued total order messages –delivered after the *ask-for-help* message. Note that the certification of the enqueued messages must wait for the recovering information to be stored in the `ws_list`, as

this structure is used in the certification process, but it is not necessary to wait to the application of these missed writesets in the database. In other words, just after the storage of the transmitted writesets in both data structures, the recovering node can act as in normal mode.

This kind of recovery inherits the main ideas of the second approach described in [13] ("Data transfer within the database system") and, up to our knowledge, had been already implemented and studied in other projects (e.g., GlobData, in order to add recovery capacity to the protocols presented in [16], but its performance results were only described in an internal project report).

4.2 Compacting

This basic procedure can be enhanced by compacting the point-to-point message in order to minimize the transmission and application time. The point-to-point recovery message has to provide all the changes in the database made from $version_i + 1$ to the current version of the recoverer node. This information can be sent in a raw mode, i.e., sending the writesets of all the transactions committed during this period of time. Then, in the recovering node, each writeset is applied in a new transaction –like any other replica does in normal function. This is the way used in the basic recovering protocol explained before.

All this procedure can be enhanced if the recoverer replica elaborates a special writeset composed by the last version of each modified object in all the transactions committed during the crash time, i.e., if the same object was modified by more than one transaction, only the last version of it would be transmitted. This special writeset would be applied in a single transaction, which can greatly improve the committing time, not only for being just one –although possibly big– transaction, but also for avoiding useless updates of the same object. This way, compacting will reduce both the transmission and the checking time as we will see later in the performance results. The time needed by the recoverer node to prepare this compacted message is not negligible, but we will see in the graphs that it does not imply any noticeable overhead.

Note also that the regular function of the replication protocol is not compromised by this optimization. Indeed, the recovering replica can start processing transactions immediately. The writesets transferred in the recovery message are not needed by the recovering replica in order to certify any new local writeset, since such new writesets should be certified against the writesets regularly delivered in the new view in which such recovering replica has rejoined the group. However, such compacted writesets can be needed for certifying remote transactions in such recovering replica, but its compacted version is enough for such kind of certification. Note that can exist long remote transactions that have started before the recovery process started, and their [start, commit] interval might overlap the end of some transactions included in the compacted missed writesets. Since at least the latest version of each missed updated item is present in such compacted set, all conflicts detectable with the original writeset list will be detectable with such compacted sequence. For instance, assume that there were N transactions T_1, T_2, ..., T_N in the original missed writeset list and that each writeset contained M items a_{11}, a_{12}, ..., a_{1M}, ..., a_{N1}, a_{N2}, ..., a_{NM}, and each of these transactions has a consecutive logical commit timestamp (t_i for T_i, being t_{i+1}=$t_i + 1$). Without generalization loss, let us assume that there are only M/2 items per writeset that have not been updated in any of its successive writesets (except in the last writeset of such compacted list that is the single one that cannot be compacted –their updated items are their trivially latest versions in the recovery transfer set–), being a_{iK_i} those items (where $K_i \subset \{j \in \mathbb{N} : 1 \leq j \leq M\}$ and $|K_i| = M/2$). So, if a given "future" transaction T_j was started between, e.g. T_1 and T_2, its writeset should be checked against all writesets in the range [T_2, T_{j-1}]. Note that T_j has been terminated after T_N, and as a result of this, all items updated by all transactions in the range [T_2, T_{j-1}] are also included in its "compacted variant" since $N < j - 1$, and our compacting process guarantees that only items rewritten during the [T_1,T_{N-1}] interval are removed from the $T_1..T_N$ writeset sequence (but if any item has not been rewritten, it appears in such compacted sequence, and this guarantees that exactly one version of all original writeset elements appears in the compacted version). Additionally, we have the advantage of a boost in the checking time, since instead of having the complete sequence of [T_1, T_N] writesets, we only have a compacted item sequence $a_{1K_1}..a_{NK_N}$, as assumed above (i.e., half of the items, in this hypothetical example).

Our optimization shares some of the characteristics of the fifth recovery strategy described in [13] ("Restricting the set of objects to check") but further optimizes that technique. To this end, our compacting is able to restrict the objects being checked without needing any additional table where the objects are being recorded during the crash interval.

Additionally, it still shares the advantage of getting such set of items to be transferred without requiring any read lock nor global read operation on the items stored in the regular database tables. But, on the other hand, it is partially dependent on the replication protocol approach (certification-based), and can not be easily adapted to all other database replication variants (e.g., the active and weak-voting variants [18] do not need any historic writeset log).

5 Performance Study

In this work we intend to measure several aspects of our recovery implementation:

- Under which circumstances (work load and crash length) a failed node can recover and reach the state of the other replicas.
- How long does it take to reach the state of the other replicas.
- Compacting impact.

To accomplish the comparison, we use PostgreSQL [15] as the underlying DBMS, and a database with a single table with two columns and 10000 rows. One column is declared as primary key, containing natural numbers from 1 to 10000 as values.

All protocols have been tested using our MADIS middleware with 4 replica nodes. Each node has an AMD Athlon(tm) 64 Processor at 2.0 GHz with 2 GB of RAM running Linux Fedora Core 5 with PostgreSQL 8.1.4 and Sun Java 1.5.0. They are interconnected by a 1 Gbit/s Ethernet. In each replica, there is a varying number of concurrent clients (from 4 to 12). Each client executes an endless stream of sequential transactions, each of one accessing a fixed number of 20 items for writing, with a fixed pause of 500 ms between each consecutive transaction. Each test begins with the execution of 500 global transactions, after that, a failure occurs in a random replica. The failure lasts for a period in which a varying number of global transactions is executed by the other replicas. After this time, the failed node restarts and begins the recovery process until it reaches the state of any of the other replicas. The test continues once the recovery ends, until the completion of 500 more global transactions, when the experiment finishes.

In the figures we show the evolution of nodes in committing transactions in the system. All the transactions have a global identifier –the end counter– and must be committed locally in each replica. This way, one global transaction requires a local transaction in each replica, and we can know how quick a node goes by seeing the last committed global identifier at that node (see the vertical axes in both Figure 2 and Figure 3). This way, each graph shows this evolution in three nodes in the system: the failed, the recoverer and another node. The bigger the slope of that curve, the faster the node goes committing global transactions.

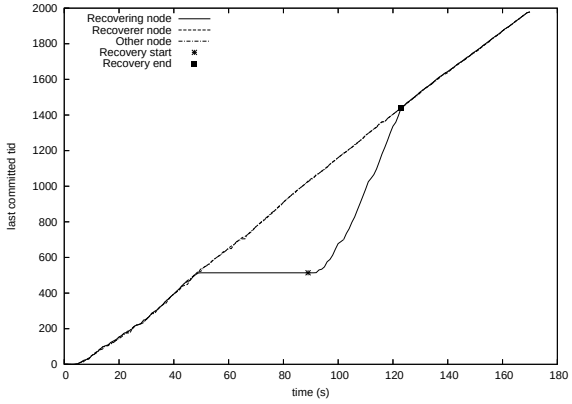
The results obtained show that the basic recovery technique was very poor in comparison with the compacting approach. This was so noticeable that we were forced to prevent new clients in the recovering node when the basic technique was used, whilst in the compacting approach such recovering replica was able to serve clients during the recovery interval.

The results obtained without compacting (see Figure 2.a) and light load show that the recovering node can easily reach the current state of the system. We can see in the figure that all the replicas have a linear evolution and when the failure occurs, the failed node does not make any advance –and so its line is horizontal. Then, when the recovery process begins, the recovering node starts to progress with more slope than the other nodes, i.e., it commits more transactions per second, and thus it can reach the global state and continue with the same previous linear behavior.

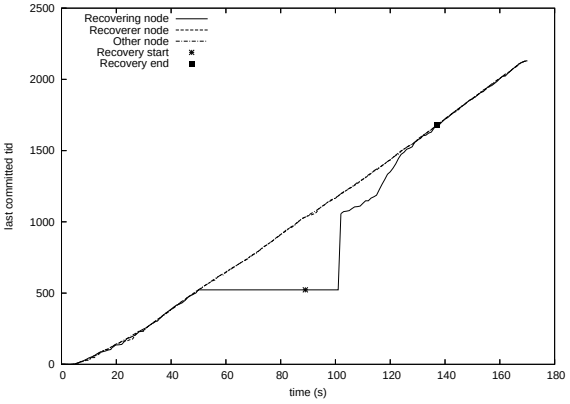
With a medium load and medium crash length (Figure 2.b), this evolution is slower and the graph is not as clear as the previous one. The previous linear behavior becomes in a curve due to the heavier load and the recovering node needs more time to recover due to the longer crash time.

Finally, with a heavy load and the maximum crash time (Figure 2.c), the behavior is similar to that of the previous case. The recovering node can reach the global state faster due to the fact that the other replicas are more loaded and, as commented before, the recovering node has no clients after the crash. This way, the recovering node can commit global transactions faster than the others and so reach the global state more quickly than when the load and the crash time are medium.

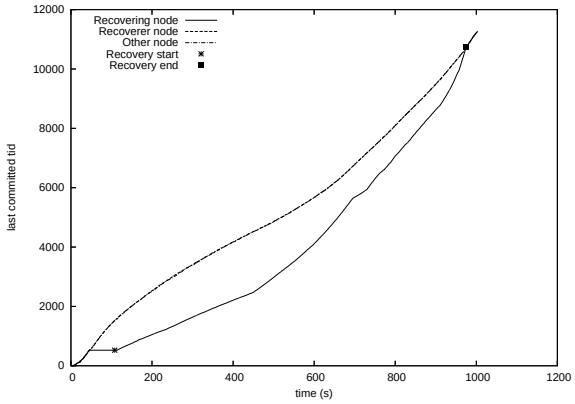
As it can be seen in the graphs, the compacting technique (see Figure 3) allows the failed replica to quickly achieve a state close to those of the other replicas. In all the tests, the recovering replica is able to do so without too much delay. Specifically, when the load and the crash time are small, we can observe that the evolution of the recovering node after the crash is not as progressive as in the basic technique, but it has two phases. The first one is a big



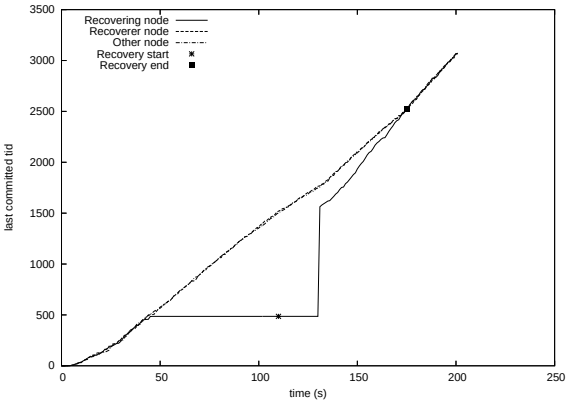
a) Light load (4 clients per node), short crash (500 transactions).



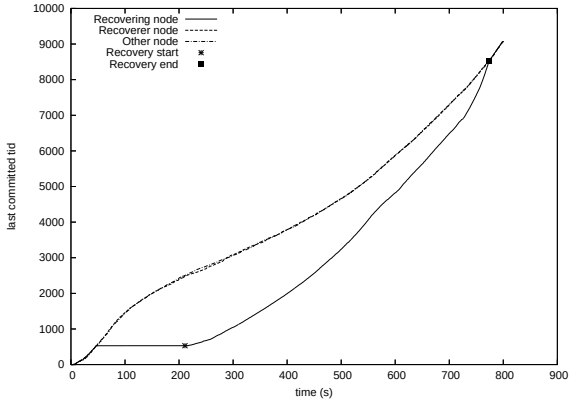
a) Light load (4 clients per node), short crash (500 transactions).



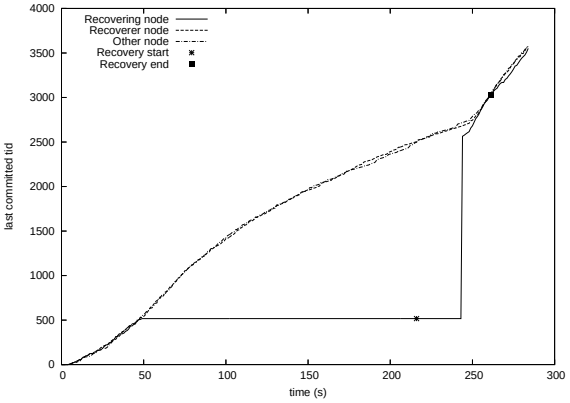
b) Medium load (8 clients per node), medium crash (1000 transactions).



b) Medium load (8 clients per node), medium crash (1000 transactions).



c) Heavy load (12 clients per node), long crash (2000 transactions).



c) Heavy load (12 clients per node), long crash (2000 transactions).

Figure 2: Recovery without compacting

Figure 3: Recovery with compacting

step towards the global state due to the application of the compacted writeset; and the second, the final evolution during the application of the enqueued messages delivered in the meanwhile. Comparing with the previous basic technique, it can be noticed that the recovery process lasts more or less like in the first case with the basic technique (i.e., with light load and a short crash interval). This shows that, in such conditions, there are no noticeable differences between the two techniques.

The next graph (Figure 3.b) shows the behavior of the system when both the load and the crash time are medium. The shape of the curves is similar to that of a light load and the recovery process is much more faster than with the basic technique. In the example provided in our figures, with the basic technique 867 seconds were required to terminate the recovery, but only 65 with the compacting optimization.

Finally, when the load and crash time are maximum (Figure 3.c), we can observe the non-linear general behavior that we mentioned in the basic technique graphs. The optimized technique still provides much better results than the basic approach (45 sec, whilst the basic one required 563 sec).

To sum up, the optimization presented in this work has been able to reduce the recovery time to a 8% of the original recovery time when 12 clients per node have been executed and 2000 transactions were missed in the crash interval. Additionally, the optimized recovery has been able to immediately serve incoming transactions in such recovering replica, whilst in our basic recovery technique this was not allowed. Moreover, with more than 2000 missed transactions and more than 12 clients, the basic recovery technique failed to complete the recovery, since the recovering replica was not able to process the queue of received writesets on time, and its receiving queue was continuously growing, whilst the optimized version did not get overloaded with twice such load.

6 Related Work

The use of version-based recovery protocols –the same approach taken as the basis of our proposed optimization– had been already suggested in the fourth and fifth recovery variants of [13], but in both cases still demanded a lot of effort for maintaining the set of versions to be transferred to each crashed replica. Either a version-based DBMS was assumed or a special additional table needs to be managed and updated each time a transaction commits. We used the latter solution in [3, 4] but in both papers such protocols were designed as a recovery approach for a replication system that did not provide any standard isolation level. Those solutions were developed in our COPLA system [8], and such middleware was targeted to provide an object-relational translation, with an object-oriented programming interface where the traditional isolation levels did not match. In the current paper, we have optimized the version-based approach taking as its basis a certification-based [18] replication variant in order to support the snapshot [2] isolation level.

Only in [3] and [11] there are some performance analyses of database recovery protocols. But, as already said, [3] is penalized by its non-standard features (non-standard API and non-standard isolation level), whilst the replication protocol assumed in [11] was hybrid (could be configured either as eager or lazy, but always with a lazy core) and this introduced a high abortion rate that was partially compensated with an outdateness estimation function. In all cases, the advantages of both approaches –and both were developed by our research group– have been improved by the solution presented now (shortest recovery time, and lowest abortion rate).

There have been many other works devoted to database replica recovery [1, 4, 10, 13] but, up to our knowledge, none of them has presented a performance study of their proposed solutions.

7 Conclusions

We have presented a first basic recovery approach for certification-based recovery protocols, analyzing its recovery time when the system load varies. Up to our knowledge this is the first performance study for such kind of recovery techniques in the field of database replication. Although certification-based replication protocols provide a good basis for developing recovery protocols, this first basic approach can be easily improved. A possible optimization based on a missed update compacting has also been presented. The performance study shows that the overall recovery time can be reduced up to a 8% of the recovery time of the basic approach (i.e., a 92% improvement), and such an improvement has been obtained allowing the immediate acceptance of incoming transactions in the recovering replica. In the basic approach this quick transaction acceptance was not possible, since it prevents the rejoining replica from completing its recovery if the system load exceeds 10 TPS in our middleware.

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