Third IEEE International Symposium on Dependable, Autonomic and Secure Computing

A Byzantine Fault Tolerant Distributed Commit Protocol [*]

Wenbing Zhao
Department of Electrical and Computer Engineering
Cleveland State University, 2121 Euclid Ave, Cleveland, OH 44115
wenbing@ieee.org

Abstract

In this paper, we present a Byzantine fault tolerant distributed commit protocol for transactions running over untrusted networks. The traditional two-phase commit protocol is enhanced by replicating the coordinator and by running a Byzantine agreement algorithm among the coordinator replicas. Our protocol can tolerate Byzantine faults at the coordinator replicas and a subset of malicious faults at the participants. A decision certificate, which includes a set of registration records and a set of votes from participants, is used to facilitate the coordinator replicas to reach a Byzantine agreement on the outcome of each transaction. The certificate also limits the ways a faulty replica can use towards non-atomic termination of transactions, or semantically incorrect transaction outcomes.

Keywords: Distributed Transaction, Two Phase Commit, Fault Tolerance, Byzantine Agreement, Web Services

1. Introduction

The two-phase commit (2PC) protocol [8] is a standard distributed commit protocol [12] for distributed transactions. The 2PC protocol is designed with the assumptions that the coordinator and the participants are subject only to benign faults, and the coordinator can be recovered quickly if it fails. Consequently, the 2PC protocol does not work if the coordinator is subject to arbitrary faults (also known as Byzantine faults [10]) because a faulty coordinator might send conflicting decisions to different participants. Unfortunately, with more and more distributed transactions running over the untrusted Internet, driven by the need for business integration and collaboration, and enabled by the latest Web-based technologies such as Web services, it is a realistic threat that cannot be ignored.

This problem is first addressed by Mohan et al. in [11] by integrating Byzantine agreement and the 2PC protocol. The basic idea is to replace the second phase of the 2PC protocol with a Byzantine agreement among the coordinator, the participants, and some redundant nodes within the root cluster (where the root coordinator resides). This prevents the coordinator from disseminating conflicting transaction outcomes to different participants without being detected. However, this approach has a number of deficiencies. First, it requires all members of the root cluster, including participants, to reach a Byzantine agreement for each transaction. This would incur very high overhead if the size of the cluster is large. Second, it does not offer Byzantine fault tolerance protection for subordinate coordinators or participants outside the root cluster. Third, it requires the participants in the root cluster to know all other participants in the same cluster, which prevents dynamic propagation of transactions. In general, only the coordinator should have the knowledge of the participants set for each transaction. These problems prevent this approach from being used in practical systems.

Rothermel *et al.* [13] addressed the challenges of ensuring atomic distributed commit in open systems where participants (which may also serve as subordinate coordinators) may be compromised. However, [13] assumes that the root coordinator is trusted, which limits its usefulness. Garcia-Molina *et al.* [6] discussed the circumstances when Byzantine agreement is needed for distributed transaction processing. Gray [7] compared the problems of distributed commit and Byzantine agreement, and provided insight on the commonality and differences between the two paradigms.

In this paper, we carefully analyze the threats to atomic commitment of distributed transactions and evaluate strategies to mitigate such threats. We choose to use a Byzantine agreement algorithm only among the coordinator replicas, which avoids the problems in [11]. An obvious candidate for the Byzantine agreement algorithm is the Byzantine fault tolerance (BFT) algorithm described in [5] because of its efficiency. However, the BFT algorithm is designed to ensure totally ordered atomic multicast for requests to a replicated stateful server. We made a number of modifications to the algorithm so that it fits the problem of atomic distributed commit. The most crucial change is made to the first phase of the BFT algorithm, where the primary coordi-

[*]This work was supported in part by Department of Energy Contract DE-FC26-06NT42853, and by a Faculty Research Development award from Cleveland State University.

nator replica is required to use a decision certificate, which is a collection of the registration records and the votes it has collected from the participants, to back its decision on a transaction's outcome. The use of such a certificate is essential to enable a correct backup coordinator replica to verify the primary's proposal. This also limits the methods that a faulty replica can use to hinder atomic distributed commit of a transaction.

We integrated our Byzantine fault tolerant distributed commit (BFTDC) protocol with Kandula, a well-known open source distributed commit framework for Web services [2]. The framework is an implementation of the Web Services Atomic Transaction Specification (WS-AT) [4]. The measurements show that our protocol incurs only moderate runtime overhead during normal operations.

2. Background

2.1. Distributed Transactions

A distributed transaction is a transaction that spans across multiple sites over a computer network. It should maintain the same ACID properties [8] as a local transaction does. One of the most interesting issues for distributed transactions is how to guarantee atomicity, *i.e.,* either all operations of the transaction succeed in which case the transaction commits, or none of the operations is carried out in which case the transaction aborts.

The middleware supporting distributed transactions is often called transaction processing monitors (or TP monitors in short). One of the main services provided by a TP monitor is a distributed commit service, which guarantees the atomic termination of distributed transactions. In general, the distributed commit service is implemented by the 2PC protocol, a standard distributed commit protocol [12].

According to the 2PC protocol, a distributed transaction is modelled to contain one coordinator and a number of participants. A distributed transaction is initiated by one of the participants, which is referred to as the initiator. The coordinator is created when the transaction is activated by the initiator. All participants are required to register with the coordinator when they get involved with the transaction. As the name suggests, the 2PC protocol commits a transaction in two phases. During the first phase (also called prepare phase), a request is disseminated by the coordinator to all participants so that they can prepare to commit the transaction. If a participant is able to commit the transaction, it prepares the transaction for commitment and responds with a "prepared" vote. Otherwise, it votes "aborted". When a participant responded with a "prepared" vote, it enters a "ready" state. Such a participant must be prepared to either commit or abort the transaction. A participant that has not sent a "prepared" vote can unilaterally abort the transaction. When the coordinator has received votes from every participant, or a pre-defined timeout has occurred, it starts the second phase by notifying the outcome of the transaction. The coordinator decides to commit a transaction only if it has received the "prepared" vote from every participant during the first phase. It aborts the transaction otherwise.

2.2. Byzantine Fault Tolerance

Byzantine fault tolerance refers to the capability of a system to tolerate Byzantine faults. It can be achieved by replicating the server and by ensuring all server replicas receive the same input in the same order. The latter means that the server replicas must reach an agreement on the input despite Byzantine faulty replicas and clients. Such an agreement is often referred to as Byzantine agreement [10].

Byzantine agreement algorithms had been too expensive to be practical until Castro and Liskov invented the BFT algorithm mentioned earlier [5]. The BFT algorithm is executed by a set of $3f + 1$ replicas to tolerate f Byzantine faulty replicas. One of the replicas is designated as the primary while the rest are backups. The normal operation of the BFT algorithm involves three phases. During the first phase (called pre-prepare phase), the primary multicasts a pre-prepare message containing the client's request, the current view and a sequence number assigned to the request to all backups. A backup verifies the request message and the ordering information. If the backup accepts the message, it multicasts to all other replicas a prepare message containing the ordering information and the digest of the request being ordered. This starts the second phase, *i.e.,* the prepare phase. A replica waits until it has collected $2f$ matching prepare messages from different replicas before it multicasts a commit message to other replicas, which starts the third phase (*i.e.,* commit phase). The commit phase ends when a replica has received $2f$ matching commit messages from other replicas. At this point, the request message has been totally ordered and it is ready to be delivered to the server application.

If the primary or the client is faulty, a Byzantine agreement on the order of a request might not be reached, in which case, a new view is initiated, triggered by a timeout on the current view. A different primary is designated in a round-robin fashion for each new view installed.

3. BFT Distributed Commit

3.1. System Models

We consider transactional client/server applications supported by an object-based TP monitor such as the WS-AT conformant framework [2] used in our implementation. For simplicity, we assume a flat distributed transaction model. We assume that for each transaction, a distinct coordinator is created. The lifespan of the coordinator is the same as the transaction it coordinates.

All transactions are started and terminated by the initiator. The initiator also propagates the transaction to other participants. The distributed commit protocol is started for a transaction when a commit/abort request is received from the initiator. The initiator is regarded as a special participant. In later discussions we do not distinguish the initiator and other participants unless it is necessary to do so.

When considering the safety of our distributed commit protocol, we use an asynchronous distributed system model. However, to ensure liveness, certain synchrony must be assumed. Similar to [5], we assume that the message transmission and processing delay has an asymptotic upper bound. This bound is dynamically explored in the adapted Byzantine agreement algorithm in that each time a view change occurs, the timeout for the new view is doubled.

We assume that the transaction coordinator runs separately from the participants, and it is replicated. For simplicity, we assume that the participants are not replicated. We assume that $3f + 1$ coordinator replicas are available, among which at most f can be faulty during a transaction. There is no limit on the number of faulty participants. Similar to [5], each coordinator replica is assigned a unique id i, where i varies from 0 to $3f$. For view v, the replica whose id i satisfies $i = v \; mod \; (3f + 1)$ would serve as the primary. The view starts from 0. For each view change, the view number is increased by one and a new primary is selected.

In this paper, we call a coordinator replica correct if it does not fail during the distributed commit for the transaction under consideration, *i.e.,* it faithfully executes according to the protocol prescribed from the start to the end. However, we call a participant correct if it is not Byzantine faulty, *i.e.,* it may be subject to typical non-malicious faults such as crash faults or performance faults.

The coordinator replicas are subject to Byzantine faults, *i.e.,* a Byzantine faulty replica can fail arbitrarily. For participants, however, only a subset of faulty behaviors are tolerated, such as a faulty participant sending conflicting votes to different coordinator replicas. Some forms of participant Byzantine behaviors cannot be addressed by the distributed commit protocol.[1]

For the initiator, we further limits its Byzantine faulty behaviors. In particular, it does not exclude any correct participant from the scope of the transaction, or include any participant that has not registered properly with the coordinator replicas, as discussed below.

To ensure atomic termination of a distributed transaction, it is essential that all correct coordinator replicas agree on the set of participants involved in the transaction. In this work, we defer the Byzantine agreement on the participants set until the distributed commit stage and combine it with that for the transaction outcome. To facilitate this optimization, we need to make the following additional assumptions.

We assume that there is proper authentication mechanism in place to prevent a Byzantine faulty process from illegally registering itself as a participant at correct coordinator replicas. Furthermore, we assume that a correct participant registers with $f + 1$ or more correct coordinator replicas before it sends a reply to the initiator when the transaction is propagated to this participant with a request coming from the initiator. If a correct participant crashes before the transaction is propagated to itself, or before it finishes registering with the coordinator replicas, either no reply is sent back to the initiator, or an exception is thrown back to the initiator. As a result, the initiator should decide to abort the transaction. The interaction pattern among the initiator, participants and the coordinator is identical to that described in the WS-AT specification [4], except that the coordinator is replicated in this work.

All messages between the coordinator and the participants are digitally signed. We assume that the coordinator replicas and the participants each has a public/secret key pair. The public keys of the participants are known to all coordinator replicas, and vice versa, while the private key is kept secret to its owner. We assume that the adversaries have limited computing power so that they cannot break the encryption and digital signatures of correct coordinator replicas.

3.2. BFTDC Protocol

Figure 1 shows the pseudo-code of the our Byzantine fault tolerant distributed commit protocol. Comparing with the 2PC protocol, there are two main differences:

– At the coordinator side, an additional phase of Byzantine agreement is needed for the coordinator replicas to reach a consensus on the outcome of the transaction, before they notify the participants.

– At the participant side, a decision (commit or abort request) from a coordinator replica is queued until at least f+1 identical decision messages have been received, unless the participant unilaterally aborts the transaction. This is to make sure that at least one of the decision messages come from a correct coordinator replica.

The distributed commit for a transaction starts when a coordinator replica receives a commit request from the initiator. If the coordinator replica receives an abort request from the initiator, it skips the first phase of the distributed commit. In any case, a Byzantine agreement is conducted on the decision regarding the transaction's outcome.

The operations of each coordinator replica is defined in the BFTDistributedCommit() method in Fig. 1. During the prepare phase, a coordinator replica sends a prepare request to every participant in the transaction. The prepare request is piggybacked with a prepare certificate, which contains the commit request sent (and signed) by the initiator.

When a participant receives a prepare request from a coordinator replica, it verifies the correctness of the signature of the message and the prepare certificate (if the participant does not know the initiator's public key, this step is

[1]For example, a Byzantine faulty participant can vote to commit a transaction while actually aborting it, and vice versa.

```
Method: BFTDistributedCommit(CommitRequest)
begin
    PrepareCert := CommitRequest;
    Append PrepareCert to PrepareRequest;
    Multicast PrepareRequest;
    VoteLog := CollectVotes();
    Add VoteLog to DecisionCert;
    decision := ByzantineAgreement(DecisionCert);
    if decision = Commit then Multicast CommitRequest;
    else Multicast AbortRequest;
    Return decision;
end
Method: PrepareTransaction(PrepareRequest)
begin
    if VerifySignature(PrepareRequest) = false then
        Discard PrepareRequest and return;
    if HasPrepareCert(PrepareRequest) = false then
        Discard PrepareRequest and return;
    if P is willing to commit T then
        Log(<Prepared T>) to stable storage;
        Send ``prepared'' to coordinator;
    else
        Log(<Abort T>); Send ``aborted'' to coordinator;
end
Method: CommitTransaction(CommitRequest)
begin
    if VerifySignature(CommitRequest) = false then
        Discard CommitRequest and return;
    Append CommitRequest to DecisionLog;
    if CanMakeDecision(commit, DecisionLog) then
        Log(<Commit T>) to stable storage;
        Send ``committed'' to coordinator;
end
Method: AbortTransaction(AbortRequest)
begin
    if VerifySignature(AbortRequest) = false then
        Discard AbortRequest and return;
    Append AbortRequest to DecisionLog;
    if CanMakeDecision(abort, DecisionLog) then
        Log(<Abort T>); Abort T locally;
        Send ``aborted'' to coordinator;
end
Method: CanMakeDecision(decision, DecisionLog)
begin
    NumOfDecisions := 0;
    foreach Message in DecisionLog do
        if GetDecision(Message) = decision then
            NumOfDecisions++;
    if NumOfDecisions >= f+1 then Return true;
    else Return false;
end
```

Figure 1. Pseudo-code for our Byzantine fault tolerant distributed commit protocol.

skipped). The prepare request is discarded if any of the verification steps fails. Even though the check for a prepare certificate is not essential to the correctness of our distributed commit protocol, it nevertheless can prevent a faulty coordinator replica from instructing some participants to prepare a transaction, even after the initiator has requested to abort the transaction.

At the end of the prepare phase, all correct coordinator replicas engage in an additional round for them to reach a Byzantine agreement on the outcome of the transaction. The Byzantine agreement algorithm used in this phase is elaborated in Section 3.3.

When a participant receives a commit request from a coordinator replica, it commits the transaction only if it has received the same decision from f other replicas so that at least one of them comes from a correct replica. The handling of an abort request is similar.

3.3. Byzantine Agreement Algorithm

The Byzantine agreement algorithm used in the BFTDC protocol is adapted from the BFT algorithm by Castro and Liskov [5]. To avoid possible confusion with the terms used to refer to the distributed commit protocol, the three phases during normal operations are referred to as ba-pre-prepare, ba-prepare, and ba-commit. Our algorithm differs from the BFT algorithm in a number of places due to different objectives. The BFT algorithm is used for server replicas to agree on the total ordering of the requests received, while our algorithm is used for the coordinator replicas to agree on the outcome (and participants set) of each transaction. In our algorithm, the ba-pre-prepare message is used to bind a decision (to commit or abort) with the transaction under concern (represented by a unique transaction id). In [5], the ba-pre-prepare message is used to bind a request with an execution order (represented by a unique sequence number). Furthermore, for distributed commit, an instance of our algorithm is created and executed for each transaction. When there are multiple concurrent transactions, multiple instances of our algorithm are running concurrently and independently from each other (the relative ordering of the distributed commit for different transactions is not important). In [5], however, a single instance of the BFT algorithm is used for *all* requests to be ordered.

When a replica completes the prepare phase of the distributed commit for a transaction, an instance of our Byzantine agreement algorithm is created. The algorithm starts with the ba-pre-prepare phase. During this phase, the primary p sends a ba-pre-prepare message including its decision certificate to all other replicas. The ba-pre-prepare message has the form $<\text{BA-PRE-PREPARE}, v, t, o, C>_{\sigma_p}$, where v is the current view number, t is the transaction id, o is the proposed transaction outcome (*i.e.,* commit or abort), C is the decision certificate, and σ_p is the signature of the message signed by the primary. The decision certificate contains a collection of records, one for each participant. The record for a participant j contains a signed registration $R_j = (t, j)_{\sigma_j}$ and a signed vote $V_j = (t, vote)_{\sigma_j}$ for the transaction t, if a vote from j has been received by the primary. The transaction id is included in each registration and vote record so that a faulty primary cannot reuse an obsolete registration or vote record to force a transaction outcome against the will of some correct participants (which may lead to non-atomic transaction commit).

A backup accepts a ba-pre-prepare message provided:

– The message is signed properly by the primary. The replica is in view v, and it is handling transaction t.

– It has not accepted a ba-pre-prepared message for transaction t in view v.

– The registration records in C are identical to, or form a superset of, the local registration records.

– Every vote record in C is properly signed by its sending participant and the transaction identifier in the record matches that of the current transaction, and the proposed decision o is consistent with the registration and vote records.

Note that a backup does not insist on receiving a decision certificate identical to its local copy. This is because a correct primary might have received a registration from a participant which the backup has not, or the primary and backups might have received different votes from some Byzantine faulty participants, or the primary might have received a vote that a backup has not received if the sending participant crashed right after it has sent its vote to the primary.

If the registration records in C form a superset of the local registration records, the backup updates its registration records and asks the primary replica for the endpoint reference[2] of each missing participant (so that it can send its notification to the participant).

A backup suspects the primary and initiates a view change immediately if the ba-pre-prepare message fails the verification. Otherwise, the backup accepts the ba-pre-prepare message. At this point, we say the replica has *ba-pre-prepared* for transaction t. It then logs the accepted ba-pre-prepare message and multicasts a ba-prepare message with the same decision o as that in the ba-pre-prepare message (this starts the ba-prepare phase). The ba-prepare message takes the form $<$BA-PREPARE$, v, t, d, o, i>_{\sigma_i}$, where d is the digest of the decision certificate C.

A coordinator replica j accepts a ba-prepare message provided:

– The message is correctly signed by replica i, and replica j is in view v and the current transaction is t;

– The decision o matches that in the ba-pre-prepare message;

– The digest d matches the digest of the decision certificate in the accepted ba-pre-prepare message.

If a replica has collected $2f$ matching ba-prepare messages from different replicas (including the replica's own ba-prepare message if it is a backup), the replica is said to have *ba-prepared* to make a decision on transaction t. This is the end of the ba-prepare phase.

A ba-prepared replica enters the ba-commit phase by multicasting a ba-commit message to all other replicas. The ba-commit message has the form $<$BA-COMMIT$, v, t, d, o, i>_{\sigma_i}$. The replica i is said to have *ba-committed*, if it has obtained $2f + 1$ matching ba-commit messages from different replicas (including the message it has sent). When a replica is ba-committed for transaction t, it sends the decision o to all participants of transaction t.

If a replica i could not advance to the *ba-committed* state until a timeout, it initiates a view change by sending a view change message to all other replicas. The view change message has the form $<$VIEW-CHANGE$, v+1, t, P, i>_{\sigma_i}$, where P contains information regarding its current state. If the replica has ba-pre-prepared t in view v, it includes a tuple $<v, t, o, C>$. If it has ba-prepared t in view v, it includes both the tuple $<v, t, o, C>$ and $2f$ matching ba-prepared messages from different replicas for t obtained in view v. If the replica has not ba-pre-prepared t, it includes its own decision certificate C.

A correct replica that has not timed out the current view multicasts a view change message only if it is in view v and it has received valid view change messages for view $v + 1$ from $f + 1$ different replicas. This is to prevent a faulty replica from inducing unnecessary view changes. A view change message is regarded as valid if it is for view $v + 1$ and the ba-pre-prepare and ba-prepare information included in P, if any, is for transaction t in a view up to v.

When the primary for view $v + 1$ receives $2f + 1$ valid view change messages for $v + 1$ (including the one it has sent or would have sent), it installs the new view, and multicasts a new view message, in the form $<$NEW-VIEW$, v + 1, V, t, o, C>$ for view $v + 1$, where V contains $2f + 1$ tuples for the view change messages received for view $v + 1$. Each tuple has the form $<i, d>$, where i is the sending replica, and d is the digest of the view change message. The proposed decision o for t and the decision certification C are determined according to the following rules:

1. If the new primary has received a view change message containing a valid ba-prepare record for t, and there is no conflicting ba-prepare record, it uses that decision.
2. Else, the new primary rebuilds a set of registration records from the received view change messages. This new set may be identical to, or a superset of, the registration set known to the new primary prior to the view change. The new primary then rebuilds a set of vote records in a similar manner. It is possible that conflicting vote records are found from the same participant (*i.e.,* , a participant sent a "prepared" vote to one coordinator replica, while sending an "aborted" vote to some other replicas), in which case, a decision has to be made on the direction of the transaction t. In this work, we choose to take the "prepared" vote to maximize the commit rate. A new decision certificate will be constructed and a decision for t's outcome is proposed accordingly. They will be included in the new view message for view $v + 1$.

When a backup receives the new view message, it verifies the message basically by following the same steps used

[2]The term endpoint reference refers to the physical contact information such as host and port of a process. In Web services, an endpoint reference typically contains a URL to a service and an identifier used by the service to locate the specific handler object [9].

by the primary. If the replica accepts the new view message, it may need to retrieve the endpoint references for some participants that it did not receive from other correct replicas. When a backup replica has accepted the new view message and obtained all missing information, it sends a ba-prepare message to all other replicas. The algorithm then proceeds according to its normal operations.

3.4. Informal Proof of Correctness

We now provide an informal proof of the safety of our Byzantine agreement algorithm and the distributed commit protocol. Due to space limitation, the proof for liveness is omitted.

Claim 1: If a correct coordinator replica ba-commits a transaction t with a commit decision, the registration records of all correct participants must have been included in the decision certificate, and all such participants must have voted to commit the transaction.

We prove by contradiction. Assume that there exists a correct participant p whose registration is left out of the decision certificate. Since a correct coordinator replica has ba-committed t with a commit decision, it must have accepted a ba-pre-prepare message and $2f$ matching ba-prepare message from different replicas. This means that a set R_1 of $2f+1$ replicas have all accepted the same decision certificate without the participant p, the initiator has requested the coordinator replicas to commit t, and every participant in the registration set has voted to commit the transaction. This further implies that the initiator has received normal replies from all participants, including p, to which it has propagated the current transaction. Because the participant p is correct and responded to the initiator's request properly, it must have registered with at least $2f+1$ coordinator replicas prior to sending its reply to the initiator. Among the $2f+1$ coordinator replicas, at least a set R_2 of $f+1$ replicas are correct, *i.e.,* all replicas in R_2 are correct and have the registration record for p prior to the start of the distributed commit for t. Because the total number of replicas is $3f+1$, the two sets R_1 and R_2 must intersect in at least one correct replica. The correct replica in the intersection either did not receive the registration from p, or it has accepted a decision certificate without the registration record for p despite the fact that it has received the registration from p, which is impossible. Therefore, all correct participants must have been included in the decision certificate if any correct replica ba-committed a transaction with a commit decision.

We next prove that if any correct replica ba-committed a transaction with a commit decision, all correct participants must have voted to commit the transaction. Again, we prove by contradiction. Assume that the above statement is not true, and a correct participant q has voted to abort the transaction t. Since we have proved above that q's registration record must have been included in the decision certificate, its vote cannot be ignored. Furthermore, since a correct replica ba-committed t with a commit decision, the set R_1 of $2f+1$ replicas have all accepted the commit decision. Again, since R_1 and R_2 must intersect by at least one correct replica, that replica both accepted the commit decision and has received the "aborted" vote from q. This is possible only if the ba-pre-prepare message that the replica has accepted contains a "prepared" vote from q. This contradict to the fact that q is a correct participant. A correct participant never sends conflicting votes to different coordinator replicas. This concludes our proof for claim 1.

Claim 2: Our Byzantine agreement algorithm ensures that all correct coordinator replicas agree on the same decision regarding the outcome of a transaction.

We prove by contradiction. Assume that two correct replicas i and j reach different decisions for t, without loss of generality, assume i decides to abort t in a view v and j decides to commit t in a view u.

First, we consider the case when $v = u$. According to our algorithm, i must have accepted a ba-pre-prepare message with an abort decision supported by a decision certificate, and $2f$ matching ba-prepare messages from different replicas, all in view v, this means a set R_3 of at least $2f+1$ replicas have ba-prepared t with an abort decision in view v. Similarly, replica j must have accepted a ba-pre-prepare message with a commit decision supported by a decision certificate, and $2f$ matching ba-prepare messages from different replicas for transaction t in the same view v, which means a set R_4 of at least $2f+1$ replicas have ba-prepared t with a commit decision in view v. Since there are only $3f+1$ replicas, the two sets R_3 and R_4 must intersect in at least $f+1$ replicas, among which, at least one is a correct replica. It means that this replica must have accepted two conflicting ba-pre-prepare messages (one to commit and the other to abort) in the same view. This contradicts the fact that it is a correct replica.

Next, we consider the case when view $u > v$. Since replica i ba-committed with an abort decision for t in view v, it must have received $2f+1$ matching ba-commit messages from different replicas (including the one sent by itself). This means that a set R_5 of $2f+1$ replicas have ba-prepared t in view v, all with the same decision to abort t. To install a new view, the primary of the new view must have received view change messages (including the one it has sent or would have sent) from a set R_6 of $2f+1$ replicas. Similar to the previous argument, the two sets R_5 and R_6 intersect in at least $f+1$ replicas, among which, at least one must be a correct replica. This replica would have included the decision and the decision certificate backed by the ba-pre-prepare message and the $2f$ matching ba-prepare messages it has received from other replicas, in its view change message. The primary in the new view, if it is correct, must have used the decision and decision certificate from this replica. This should have led all correct replicas to ba-commit transaction t with an abort decision, which contradicts to the assumption that a correct replica committed t. If the primary is faulty and did not obey the new view construction rule, we argue that no correct replica could have accepted the new view message, let alone to have ba-committed t with a commit decision. Recall that a correct replica should verify the new view message by following the new view construction rules, just as a correct primary would do. We have proved above that the $2f + 1$ view change messages must contain one sent by a correct replica with ba-prepare information for an abort decision. A correct replica cannot possibly have accepted the new view message sent by the faulty primary, which contains a conflicting decision. This contradicts to the initial assumption that a correct replica j committed transaction t in view u. The proof for the case when $v > u$ is similar. Therefore, claim 2 is correct.

Claim 3: The BFTDC protocol guarantees atomic termination of transactions at all correct participants.

We prove by contradiction. Assume that a transaction t commits at a participant p but aborts at another participant q. According to the criteria indicated in the CommitTransaction() method shown in Fig. 1, p commits the transaction t only if it has received the commit request from at least $f+1$ different coordinator replicas. Since at most f replicas are faulty, at least one request comes from a correct replica. Due to claim 1, if any correct replica ba-committed a transaction with a commit decision, then the registration records of all correct participants must have been included in the decision certificate, and all correct participants must have voted to commit the transaction.

On the other hand, since q aborted t, one of the following two scenarios must be true: (1) q unilaterally aborted t, in which case, it must not have sent a "prepared" vote to any coordinator replica. (2) q received a prepare request, prepared t, sent a "prepared" vote to one or more coordinator replicas. But it received an abort request from at least $f + 1$ different coordinator replicas.

If the first scenario is true, q might or might not have finished its registration process. If it did not, the initiator would have been notified by an exception, or would have timed out q. In any case, the initiator should have decided to abort t. This conflicts with the fact that p has committed t because it implies that the initiator has asked the coordinator replicas to commit t. If q completed the registration process, its registration record should have been aware by a set R_7 of at least $f+1$ correct replicas. Since p has committed t, at least one correct replica has ba-committed t with a commit decision, which in turn implies that a set R_8 of at least $2f + 1$ coordinator replicas have accepted a ba-pre-prepare message with a decision certificate either has no q in its registration records, or without q's "prepared" vote. Since there are $3f + 1$ replicas, R_7 and R_8 must intersect in at least one replica. This correct replica could not possibly have accepted a ba-pre-prepare message with a decision certificate described above.

For the second scenario, at least one correct replica has decided to abort t. Since another participant p committed t, at least one correct replica has decided to commit t. This contradicts to claim 2, which we have proved to be true. Therefore, claim 3 is correct.

4. Implementation and Performance

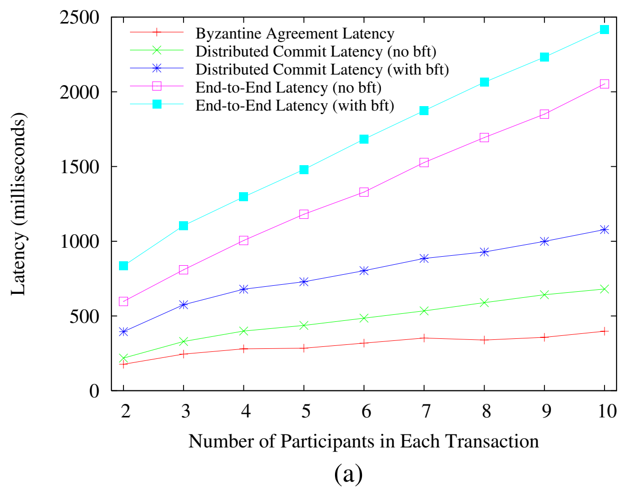
We have implemented the BFTDC protocol (with the exception of the view change mechanisms) and integrated it into a distributed commit framework for Web services in Java programming language. The extended framework is based on a number of Apache Web services projects, including Kandula (an implementation of WS-AT) [2], WSS4J (an implementation of the Web Services Security Specification) [3], and Apache Axis (SOAP Engine) [1]. Most of the mechanisms are implemented in terms of Axis handlers that can be plugged into the framework without affecting other components. Some of the Kandula code is modified to enable the control of its internal state, to enable a Byzantine agreement on the transaction outcome, and to enable voting. Due to space constraint, the implementation details are omitted.

For performance evaluation, we focus on assessing the runtime overhead of our BFTDC protocol during normal operations. Our experiment is carried out on a testbed consisting of 20 Dell SC1420 servers connected by a 100Mbps Ethernet. Each server is equipped with two Intel Xeon 2.8GHz processors and 1GB memory running SuSE 10.2 Linux.

The test application is a simple banking Web services application where a bank manager (*i.e.,* initiator) transfers funds among the participants within the scope of a distributed transaction for each request received from a client. The coordinator-side services are replicated on 4 nodes to tolerate a single Byzantine faulty replica. The initiator and other participants are not replicated, and run on distinct nodes. The clients are distributed evenly (whenever possible) among the remaining nodes. Each client invokes a fund transfer operation on the banking Web service within a loop without any "think" time between two consecutive calls. In each run, 1000 samples are obtained. The end-to-end latency for the fund transfer operation is measured at the client. The latency for the distributed commit and the Byzantine agreement is measured at the coordinator replicas. Finally, the throughput of the distributed commit framework is measured at the initiator for various number of participants and concurrent clients.

To evaluate the runtime overhead of our protocol, we compare the performance of our BFTDC protocol with the 2PC protocol as it is implemented in the WS-AT framework with the exception that all messages exchanged over the network are digitally signed.

In Figure 2(a), we included the distributed commit latency and the end-to-end latency for both our protocol (indicated by "with bft") and the original 2PC protocol (indicated by "no bft"). The Byzantine agreement latency is also shown. Figure 2(b) shows the throughput measurement re-



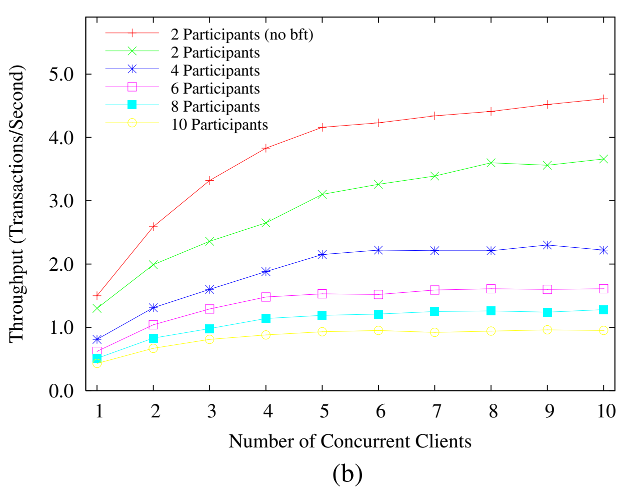


Figure 2. (a) Various latency measurements for transactions with different number of participants under normal operations (with a single client). (b) Throughput of the distributed commit service in terms of transactions per second for transactions with different number of participants under different load.

sults for transactions using our protocol with up to 10 concurrently running clients and 2-10 participants in each transaction. For comparison, the throughput for transactions using the 2PC protocol for 2 participants is also included.

As can be seen in Figure 2(a), the latency for the distributed commit and the end-to-end latency both are increased by about 200-400 ms when the number of participants varies from 2 to 10. This increase is mostly attributed to the introduction of the Byzantine agreement phase in our protocol. Percentage-wise, the end-to-end latency, as perceived by an end user, is increased by only 20% to 30%, which is quite moderate. We observe a similar range of throughput reductions for transactions using our protocol, as shown in Figure 2(b).

5. Conclusions

In this paper, we presented a Byzantine fault tolerant distributed commit protocol. We carefully studied the types of Byzantine faults that might occur to a distributed transactional systems and identified the subset of faults that a distributed commit protocol can handle. We adapted Castro and Liskov's BFT algorithm to ensure Byzantine agreement on the outcome of transactions. We also proved informally the correctness of our BFTDC protocol. A working prototype of the protocol is built on top of an open source distributed commit framework for Web services. The measurement results of our protocol show only moderate runtime overhead. We are currently working on the implementation of the view change mechanisms and exploring additional mechanisms to protect a TP monitor against Byzantine faults, not only for distributed commit, but for activation, registration, and transaction propagation as well.

References

[1] Apache Axis project. http://ws.apache.org/axis/.
[2] Apache Kandula project. http://ws.apache.org/kandula/.
[3] Apache WSS4J project. http://ws.apache.org/wss4j/.
[4] L. Cabrera et al. *WS-AtomicTransaction Specification*, August 2005.
[5] M. Castro and B. Liskov. Practical Byzantine fault tolerance and proactive recovery. *ACM Transactions on Computer Systems*, 20(4):398–461, November 2002.
[6] H. Garcia-Molina, F. Pittelli, and S. Davidson. Applications of Byzantine agreement in database systems. *ACM Transactions on Database Systems*, 11(1):27–47, March 1986.
[7] J. Gray. A comparison of the Byzantine agreement problem and the transaction commit problem. *Springer Verlag Lecture Notes in Computer Science*, 448:10–17, 1990.
[8] J. Gray and A. Reuter. *Transaction Processing: Concepts and Techniques*. Morgan Kaufmann Publishers, San Mateo, CA, 1983.
[9] M. Gudgin and M. Hadley. *Web Services Addressing 1.0 - Core*. W3C working draft, February 2005.
[10] L. Lamport, R. Shostak, and M. Pease. The Byzantine generals problem. *ACM Transactions on Programming Languages and Systems*, 4(3):382–401, July 1982.
[11] C. Mohan, R. Strong, and S. Finkelstein. Method for distributed transaction commit and recovery using Byzantine agreement within clusters of processors. In *Proceedings of the ACM symposium on Principles of Distributed Computing*, pages 89–103, Montreal, Quebec, Canada, 1983.
[12] The Open Group. *Distributed Transaction Processing: The XA Specification*, February 1992.
[13] K. Rothermel and S. Pappe. Open commit protocols tolerating commission failures. *ACM Transactions on Database Systems*, 18(2):289–332, June 1993.