SYSTEMATIC DEVELOPMENT OF A FAMILY OF FAIR EXCHANGE PROTOCOLS

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Abstract: Exchange protocols play an important role in application areas such as e-commerce where protocol participants require mutual guarantees that a transaction involving exchange of items has taken place in a specific manner. A protocol is fair if no protocol participant can gain any advantage over an honest participant by misbehaving. This paper presents a family of fair exchange protocols for two participants under a variety of assumptions concerning participant misbehaviour, node reliability and message delays.

Keywords: Fair Exchange, Security, Digital Signatures, Crash tolerance, Distributed Systems, Trusted Third Party.

1. INTRODUCTION

Fair exchange protocols play an important role in application areas where protocol participants require mutual guarantees that an exchange of data items has taken place in a specific manner. An exchange is fair if a dishonest participant cannot gain any advantage over honest participants by misbehaving. Two-participant fair-exchange protocols that make use of a trusted third party have been studied extensively in the literature (e.g., [1, 5, 11]); these protocols maintain fairness even if the dishonest participant can tamper with the protocol execution in an unrestricted (malicious) manner. They however require that an honest participant's node execute the protocol correctly – suffering no failures. We call them non-fault-tolerant. A crash-tolerant fair exchange protocol, on the other hand, ensures no loss of fairness to an honest participant even if the participant's node crashes during the execution, provided it eventually recovers [4].

In this paper we develop a number of fair exchange protocols under a variety of assumptions concerning i) user misbehaviour, ii) node failures and iii) communication delays. The development is systematic: we begin by classifying dishonest participants into *restricted abusers* (they cannot tamper

with the protocol execution in an arbitrary manner) and unrestricted or *malicious abusers*, and the communication model into *synchronous*, where a known bound on message delays exists, and *asynchronous*; we develop the first protocol under the most constrained set of assumptions: restricted abuser, synchronous communication and no fault tolerance. We then relax the restricted abuser assumption to malicious abuser and then the synchrony assumption into asynchrony. The resulting family of non-fault-tolerant protocols is then transformed into a family of crash-tolerant ones.

A major advantage of our approach is that it enables a reader to gain a deeper understanding of the impact of a given set of assumptions on the problem of fair exchange; it also highlights the relationships that exist between fairness, fault tolerance and communication delay assumptions. We first describe the problem and the underlying system models (section 2), and then develop a family of non-fault-tolerant protocols (section 3), followed by their crash-tolerant counterparts (section 4). Section 5 summarises the characteristics of the protocols developed and concludes the paper.

2. SYSTEM MODELS AND THE PROBLEM DESCRIPTION

We consider two mutually untrusting users, U_A and U_B , who have data items I_A and I_B respectively that are unknown to the other user prior to the exchange. User U_x , $X \in \{A, B\}$, intends to send I_x in return for receiving I_y , where $Y \in \{A, B\}$ and $Y \neq X$. P_x denotes the process that executes an exchange protocol on behalf of user U_x on his node N_x .

Figure 1. The Fair-Exchange System.

Our distributed exchange system (Figure 1) has a third node hosting the trusted third party (TTP) process. The exchange must preserve fairness as

well as *non-repudiation*: a user gets irrefutable evidence of the actions performed by his process (to guard against the other user's false denial of the occurrence of an action or event).

2.1. Classifying User Misbehaviour

Generally, the problem of fair exchange is addressed in a context where a dishonest user, say U_B , *totally controls* the behaviour of P_B and tries to undermine every attempt to ensure fairness and non-repudiation. We term those dishonest users as *malicious abusers* and distinguish such users from a class of *restricted abusers*: a dishonest user is said to be a restricted abuser if he is unable to obtain the values of variables or procedures contained in a piece of (binary) code which his local process receives from the TTP and uses during protocol execution. Say, user U_B is a restricted abuser and the TTP gives P_B a piece of code Π_B that contains, say, some secret encryption keys. By our definition, U_B cannot obtain the keys in Π_B by visual inspection or through a trace analysis even after P_B has executed U_B . The definition does not however restrict the ability of U_B to delay, block, inspect, or tamper with any message which an execution of Π_B generates or is destined to receive.

One way to realise our notion of a restricted abuser is to have each P_X receive the TTP code Π_X via a secure channel and execute it in a tamper-proof part of the user node. We note here that there is much interest in developing tamper-proof computing subsystems by the industry led Trusted Computing Platform Alliance, TCPA [8]. Further, smartcards can be used to achieve such tamper-proof, computing subsystems for small, critical computations [2, 10]. Thus, developing fair-exchange protocols under the restricted abuser model is of practical interest, and also exposes the benefits that the model has to offer.

2.2. Classifying Node Behaviour

A fair-exchange protocol is fault-tolerant if it ensures that fairness is preserved to an honest participant despite the participant's node suffers failures of the assumed type. In other words, an honest user does not suffer loss of fairness when his node is unreliable. We consider two types of node behaviour:

Reliable: An honest user's node never fails.

Crash-recovery: An honest user's node fails by stopping to function (crash); it recovers within some finite (but unknown) time after a crash and

may crash again after its recovery. An honest user's node has access to a stable store whose contents survive the node crash.

When a user node is regarded to be reliable, the cause of any node crash is attributed to the user who has misbehaved and is therefore not entitled to any fairness guarantee. This is the traditional view taken in the TTP-based protocols (e.g., [1, 5, 11]) which are therefore not fault-tolerant (see [4]). In line with the existing, non-fault-tolerant protocols, the TTP is assumed to be *reliable* and *secure* against intrusions and Trojan horses.

2.3. Classifying Communication Delays

Communication between correctly functioning processes is assumed to be resilient to network failures: a message loss is tolerated by retransmission and a message corruption is detected using encryption and reduced to a message loss. A network intruder is therefore assumed to be a restricted abuser: a message in transit is a black box to him; he cannot modify it and have it undetected at the destination. He can only delay a message from being accepted at a destination. We consider two types of network intrusion: intruder gives up delaying a given message transmission after a known period of time or after some unknown finite time, leading to two models: in the *synchronous* model, correct processes exchange messages with delays bounded by a known D ; in the *asynchronous* model, D is finite but unknown.

2.4. Properties of a fair exchange Protocol

A user is *honest* if he makes no attempt to modify the behaviour of a protocol process except through the permitted operations.

Termination: Any execution of the protocol terminates for an honest user U_X . Termination for an honest U_X can be either a *normal* termination in which P_X delivers I_Y to U_X or an *exceptional* termination where P_X informs U_X that the exchange is unsuccessful.

No_Loss_Of_Fairness: If P_X of honest U_X terminates exceptionally, U_Y does not receive I_X .

Non-repudiation: when P_X delivers I_Y to honest U_X , it also provides irrefutable evidence against any false denial on I_Y having been sent by U_Y .

The above properties are met trivially if P_A and P_B terminate always exceptionally.

Non-Triviality: If U_A and U_B are honest, both are guaranteed to have normal termination.

Non-triviality is easy to guarantee for the synchronous, non-fault-tolerant case. When the fair-exchange protocol is deterministic (not probabilistic) in

nature, honesty of both the users alone is not sufficient to guarantee non-triviality for the case of crash-recovery nodes and synchronous communication, as well for the case of asynchronous communication. The reasons for this inability are as follows. In crash-recovery model, the bound on the time taken by an honest user node to recover from a crash is unknown and a dishonest user may never allow his node to recover. So, a correctly functioning process (of an honest user or the TTP) that is waiting too long for a message from a user process cannot resolve whether the latter is honest and its message is unduly delayed due to a crash or is dishonest and is not going to transmit the expected message; similarly, in the asynchronous model, it cannot resolve whether the user process from whom a message is expected is honest and its message is still in transit or is dishonest and the message will not be transmitted at all. (Readers familiar with the FLP impossibility result [3] may find these arguments remarkably similar.) Therefore, meeting the termination property would mean that an execution of the protocol may have to terminate exceptionally even if both users are honest. Therefore, the necessary condition for ensuring non-triviality is that the honest users be willing to re-execute the protocol after every exceptional termination; the sufficient condition differs with the combination of assumptions and is stated below: there must be an execution in which

- (i) user nodes do not crash, for the case of crash-recovery nodes and synchronous communication;

- (ii) message delays do not increase, for the case of reliable nodes and asynchronous communication; and,

- (i) and (ii), for the case of crash-recovery nodes and asynchronous communication.

2.5. Assumptions, Notations and the Exchange Preliminaries

In all our protocols, the TTP, assumed to be secure and reliable, is used to set-up the context for, and to initiate, the actual exchange. We make the following additional assumptions.

A1: Processing within correctly functioning nodes is synchronous: delays for task scheduling and processing are bounded by a known constant.

A2: The clocks of the TTP and the functioning nodes of honest users are synchronised to real-time within a known bound which, for simplicity, is assumed to be zero. A recovering node receives the current time from its user.

Notations

M: a message;

V_A (V_B): procedure to verify if I_A (I_B) satisfies its description Σ_A (Σ_B) advertised by U_A (U_B).

$eK(M)$: encryption of M using key K;

$Sig_X(M)$: signature of P_X , $X \in \{A, B\}$, on M using the private key of U_X ;

$Sig_{TTP}(M)$: TTP's signature on M. Signatures are both signer- and content-dependent, and are verifiable using the signer's public key.

L: a label that identifies a given exchange;

N: a large random number (*nonce*) generated securely by TTP to uniquely identify messages of a given execution of the protocol for a given exchange; and,

H: one-way and collision-resistant hash function: it is not feasible to compute from $H(N)$, N nor another N' such that $H(N) = H(N')$.

Π_x : contents of a message sent by the TTP to P_x to initiate the exchange phase.

When a party $Z \in \{A, B, TTP\}$ sends a message M it also includes $Sig_Z(M)$ as the evidence of origin for M. A recipient accepts a received M only if the accompanying $Sig_Z(M)$ is found authentic. For brevity, we will not make the details of this verification explicit and a received message will only refer to a message received with an authentic evidence of origin; further the pair $\{M, Sig_Z(M)\}$ will simply be written as M which will be indicated by its significant fields.

Exchange Preliminaries

All our protocols are structured into two phases: the *start-up* phase followed by the *exchange phase* when the actual exchange of promised data items takes place. During the start-up phase, the following three steps are carried out:

Step 0.1: Each U_X approaches a trusted authority TA_X to generate a verification procedure V_X using which any party can verify if a data item I satisfies specification Σ_X . For example, if I_B is a piece of software S, U_B must approach a software licensing authority (trusted by U_A) who evaluates S against the specification Σ_B that U_B advertised. If satisfied, the authority computes an evidence of evaluation $E_B = H(S)$; V_B is then generated as a program which contains E_B and evaluates the predicate $E_B = H(I)$ when invoked to verify a data item I. Using the conventional notations, the step is:

$U_A \rightarrow TA_A: (I_A, \Sigma_A);$ $TA_A \rightarrow U_A: \{V_A, \Sigma_A, Sig_{TA_A}(V_A, \Sigma_A)\};$	$U_B \rightarrow TA_B: (I_B, \Sigma_B);$ $TA_B \rightarrow U_B: \{V_B, \Sigma_B, Sig_{TA_B}(V_B, \Sigma_B)\};$
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Step 0.2: Users exchange their $\{V, \Sigma, \text{Sig}_{TA}(V, \Sigma)\}$, decide on the TTP and on a (future) time T_{EX} when the TTP should initiate the exchange phase.

Step 0.3: $U_A \rightarrow TTP: (A, B, T_{EX}, \{V_A, \Sigma_A\}, \{V_B, \Sigma_B\}, \text{rtt}_{AB})$; similarly, $U_B \rightarrow TTP: (A, B, T_{EX}, \{V_A, \Sigma_A\}, \{V_B, \Sigma_B\}, \text{rtt}_{BA})$. The last field rtt_{XY} is U_X 's estimation of the round trip time between its node and N_Y .

Upon receiving messages sent in step 0.3, the TTP verifies whether all fields except the last one are identical. If so, the TTP decides on a unique L for the exchange if the exchange is new; generates a nonce N for the exchange phase and records N in a variable $active_run\#(L) = N$. If the exchange is not new, it only generates a new nonce and records it in $active_run\#(L)$. It then sends Π_B to N_A and Π_B to N_B which include the parameters L and $H(N)$ which are used to identify messages of a given execution. The nature and the complete contents of Π 's sent by TTP depend on the exchange phase of the protocol for a chosen combination of various assumptions discussed earlier.

TTP Involvement

A protocol is said to keep TTP *offline* if it is possible for honest user processes to achieve normal termination without interacting with TTP after the exchange phase is initiated. Such a protocol is also called *optimistic* in the literature. If TTP is not off-line, it is said to be *on-line*.

A protocol is said to use a *state-keeping* TTP if it requires the TTP to respond to messages from user processes for an unspecified amount of time; the protocol is said to use a *state-relinquishing* TTP if it requires the TTP only for a specified duration after the exchange phase has been initiated. From cost point of view, TTP is preferred to be offline and state-relinquishing.

3. NON-FAULT-TOLERANT FAIR EXCHANGE PROTOCOLS

3.1. Restricted Abuser, Synchronous Communication (Protocol P1)

The TTP sends a code Π_A to P_A and Π_B to P_B . P_A and P_B perform the exchange phase of the protocol P1 by executing the code given to them. Embedded in Π_X are L , $H(N)$, V_X , Δ , and keys K_A and K_B , each with TTP's evidence of origin, Δ is set to D (the known bound on message delays), K_A and K_B are symmetric and random session keys. When a dishonest U_X is only a restricted abuser, he cannot obtain K_A and K_B from Π_X nor undetectably modify V_X embedded within Π_X . However, U_X can delay, block, inspect, or

tamper with any message P_X generates or is destined to receive while it executes Π_X .

On executing Π_A , P_A first verifies whether I_A (input by U_A) passes V_A embedded within Π_A . Only if I_A passes the verification, P_A continues the execution: it encrypts the verified I_A using K_A , forms a message M_A and encrypts M_A with K_B . The encrypted M_A , like all messages sent by P_A , contains a signature generated using the private key of U_A and is sent to P_B . (Any message without valid signature is ignored and hence assumed not 'received' by a destination.) P_B decrypts the received M_A , and if the contents are correct, it sends an acknowledgement $Ack_B(A)$ to P_A . P_B similarly sends M_B to, and expects to receive $Ack_A(B)$ from, P_A . The message exchange between P_A and P_B are depicted in Figure 2, where a process sending a message shown along an out-going arrow is conditional upon that process having received the messages shown along all incoming arrows. P_X , after receiving Π_X , starts executing both the rounds concurrently: send M_X and wait for $Ack_Y(X)$ while waiting for M_Y before $Ack_X(Y)$ can be sent; a timeout of 2Δ is used to avoid indefinite waiting.

Figure 2. Protocol P1: Message exchange rounds between user processes.

Table 1 summarises the messages used and the intended direction of their flow. A message is indicated by its significant fields, the first three fields will be: the label L , the sender, and $H(N)$. Note that the first and the third message fields uniquely refer to a given execution.

The condition on P_A (P_B) for computing I_B (I_A) from M_B (M_A) is: both M_B (M_A) and $Ack_B(A)$ ($Ack_A(B)$) must be received within 2Δ time after receiving Π_A (Π_B) from TTP. If P_A receives only M_B (within 2Δ time), it asks TTP to *resolve* the exchange by sending message **Res_A** that contains both M_A and the received M_B . If it receives neither M_B nor $Ack_B(A)$, it *requests* TTP to abort the exchange by sending message **Req_A** that contains only M_A . TTP, having initiated the exchange at its clock time T_{EX} , executes the following steps at $T_{EX} + 4\Delta$:

Step T1: if it has received **Req** from both processes or **Res** from at least one process, it sets variable *outcome* = *resolved* and resolves the exchange by sending $M_{TTP}(X)$ to both processes;

Step T2: if it has received Req from only one process, it sets *outcome* = *aborted* and aborts the exchange by sending **Abort_{TTP}(X)** to both processes;

Step T3: if it has received no message, it sets variable *outcome* to *unknown*;

Step T4: it terminates (as a state-relinquishing TTP).

$\Phi_A = eK_A(I_A)$	$\Phi_B = eK_B(I_B)$
$M_A = eK_B(L, A, H(N), \Phi_A); A \rightarrow B$	$M_B = eK_A(L, B, H(N), \Phi_B); B \rightarrow A$
$Ack_A(B) = (L, A, H(N), H(M_B), My_ack); A \rightarrow B$	$Ack_B(A) = (L, B, H(N), H(M_A), My_ack); B \rightarrow A$
$Res_A = (L, A, H(N), M_A, M_B, Resolve_request); A \rightarrow TTP$	$Res_B = (L, B, H(N), M_B, M_A, Resolve_request); B \rightarrow TTP$
$Req_A = (L, A, H(N), M_A, Abort_request); A \rightarrow TTP$	$Req_B = (L, B, H(N), M_B, Abort_request); B \rightarrow TTP$
$M_{TTP}(A) = (L, TTP, H(N), M_B, My_ack); TTP \rightarrow A$	$M_{TTP}(B) = (L, TTP, H(N), M_A, My_ack); TTP \rightarrow B$
$Abort_{TTP}(A) = (L, TTP, H(N), Abort_granted, A); TTP \rightarrow A$	$Abort_{TTP}(B) = (L, TTP, H(N), Abort_granted, B); TTP \rightarrow B$

Table 1. Description of messages used in Protocol P1.

An honest P_A terminates normally by receiving either M_B and $Ack_B(A)$ or $M_{TTP}(A)$; exceptionally by receiving $Abort_{TTP}(A)$. The protocol keeps TTP off-line (step T3 above). The four properties of fair-exchange (Section 2.4) are met by P1 because: (i) dishonest user, say U_B , is a restricted abuser, so he cannot force P_B to deliver I_A against the protocol conditions nor can he obtain K_A and K_B from Π_B ; (ii) an honest P_A sends $Ack_A(B)$ to P_B only if it receives M_B in a timely manner; (iii) any message that P_A sends to TTP reaches before $T_{EX}+4\Delta$; and, (iv) TTP's response, if any, is identical to both P_A and P_B .

3.2. Malicious Abuser, Synchronous Communication (P1.1)

1. P1.1 is the same as P1 except some messages of the exchange phase have different contents. This is because a dishonest user, say U_B , can obtain the keys K_A and K_B in the program Π_B supplied by TTP at the start of the exchange phase. The following changes are needed:

2. Π_A is a message containing all parameters as before except K_B : $\Pi_A = (L, H(N), V_A, \Delta, K_A)$; similarly, $\Pi_B = (L, H(N), V_B, \Delta, K_B)$. Further, TTP should not have used K_A and K_B before.
3. Π_A does not contain K_B means that P_A cannot encrypt M_A with K_B (see Table 1). Therefore M_A is $(L, A, H(N), \Phi_A)$ and M_B is $(L, A, H(N), \Phi_B)$.
4. P_A includes K_A in its $Ack_A(B)$ so that if P_B has both M_A and $Ack_A(B)$ it can terminate normally without having to contact TTP. So, $Ack_A(B)$ of Table 1 becomes $(L, A, H(N), H(M_B), K_A)$ and $Ack_B(A)$ becomes $(L, B, H(N), H(M_A), K_B)$.
5. Finally, $M_{TTP}(A)$ by which TTP instructs P_A to resolve the exchange, must contain K_B : $M_{TTP}(A) = (L, TTP, H(N), M_B, K_B)$. Also, $M_{TTP}(B) = (L, TTP, H(N), M_A, K_A)$.

Remark: *Dispute Resolution:* Non-repudiation guarantee ensures that when honest P_A terminates normally, it has evidence that P_B sent the item which P_A delivered to U_A . However, if U_B is dishonest, U_A might find I_B not passing the verification test V_B which it agreed with U_B in the start-up phase. Consider this scenario: malicious U_B obtains K_B from Π_B , generates M'_B with authentic evidence of origin but with $I'_B \neq I_B$, and transmits M'_B in place of M_B . Since Π_A does not contain K_B (see modification 1 above), P_A cannot check whether the I'_B in the received M'_B meets V_B at the end of the first round itself (see figure 2). Since M'_B is found to have authentic evidence of origin, it will send $Ack_A(B)$ that contains K_A , thus letting P_B terminate normally without ever contacting TTP. Only after being delivered of I'_B , U_A can find out that I'_B does not pass V_B and that U_B has been dishonest.

If U_B is a restricted abuser, the above scenario cannot arise: Π_B is tamper-proof relative to the abilities of U_B and therefore it can contain K_A using which P_B encrypts M_B in P1. Since U_B cannot obtain K_A , he cannot modify M_B to M'_B and make P_A accept M'_B . Thus, in P1, what U_A receives is guaranteed to be the expected one.

P1.1 only guarantees that P_A delivers to an honest U_A what a malicious U_B *actually* sent in exchange for I_A , not necessarily what U_B has pledged to send. The TTP can guarantee fairness to U_A in such a scenario, by contacting the trusted agent TA_B employed to generate V_B which U_A accepted. This corresponds to a weaker form of fairness enforcement in the hierarchy of [9]: fairness can only be guaranteed with the help of a trusted agent outside the fair-exchange system (figure 1) and without any cooperation from U_B . The issue of post-exchange dispute resolution is discussed and addressed in [6, 7] for malicious abusers. The protocols of [6, 7] make use of *inverse* and *compatible* keys to verify an encrypted item without decrypting it. Using these special types of encryption keys, P1.1 can also be modified to ensure that an honest user receives in a normal termination only the item he bargained for. The modification is simple and shown in [2]. It preserves the

2-round structure which is provably optimal [5]. The protocol of [6], like P1.1, keeps the TTP off-line but requires 4 rounds when both users are honest.

3.3. Restricted and Malicious Abusers, Asynchronous Communication (P1A and P1.1A)

Protocols P1 and P1.1 are not appropriate for the asynchronous model. This is not surprising as [5] shows that, in the non-fault-tolerant case, there cannot be an asynchronous protocol that works with a TTP that is off-line *and* state-relinquishing, and goes on to present a protocol with an off-line and state-keeping TTP. We here present protocols P1 A and P1.1A with an on-line and state-relinquishing TTP, by deriving them from P1 and P1.1 respectively. The derivation is simple: in round 2 (see figure 2), P_X sends $Ack_X(Y)$ to TTP, not to P_Y ; it then waits for either $M_{TTP}(X)$ or $Abort_{TTP}(X)$.

In the start-up phase, TTP should obtain round trip time (rtt) measurements with each P_X (as rtt_A and rtt_B) and estimate 2Δ maximum $\{rtt_A, rtt_B, rtt_{AB}, rtt_{BA}\}$.

TTP:

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when (clock =  $T_{EX}$ ) do {
  send  $\Pi_A$  to  $P_A$ ; send  $\Pi_B$  to  $P_B$ ; Set_of_M  $M\_Bag_L = \{ \}$ ;
  repeat {receive(M); deposit M in  $M\_Bag_L$ ;} until clock <  $T_{EX} + 4\Delta$ ;
  if ( $Ack_A(B) \in M\_Bag_L$  and  $Ack_B(A) \in M\_Bag_L$ )
    then {send  $Ack_B(A)$  to  $P_A$ ; send  $Ack_A(B)$  to  $P_B$ ;} // resolved
    else {send  $Abort_{TTP}(A)$  to  $P_A$ ; send  $Abort_{TTP}(B)$  to  $P_B$ ;} } /* end do

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Figure 3. Pseudo-code for TTP in protocols P1A and P1.1 A.

If honest users re-execute the protocol for a given exchange after every exceptional termination, *non-triviality* is guaranteed if there exists an execution in which the message transfer delays between P_A , P_B and TTP do not exceed the Δ determined by TTP at the start of the exchange phase; i.e., if message transfer delays do not increase during an execution.

4. CRASH-TOLERANT PROTOCOLS

4.1. Restricted Abuser, Synchronous Communication (P2)

Two extensions are necessary for protocol P1 to be crash tolerant. *Pessimistic (synchronous) checkpointing*: P_X of an honest user logs every received message and check-points its state before the received message is processed. Thus, in fig. 2, an honest P_X must receive *and* log the messages of the incoming channels before it can send out a message. *State-keeping TTP*:

in step T4, the TTP does not terminate the execution, but continues to operate depending on the value of *outcome*: if *outcome* is *unknown*, the TTP executes steps T1 – T3 after a period of 4Δ time; otherwise, it responds to a message from P_X with $M_{TTP}(X)$ or $Abort_{TTP}(X)$, if *outcome* is *resolved* or *aborted* respectively. This modification permits a recovered P_X to terminate. Note that when TTP sets *outcome* to *resolved* or *aborted*, it is irreversible.

Impossibility of making P1 Crash-tolerant

Even with the use of pessimistic logging and state-keeping TTP, P1 cannot become crash-tolerant even if the dishonest user is a restricted abuser. Consider the following scenario. Let N_A crash during the second round shown in figure 2 and before P_A receives $Ack_B(A)$; that is, P_A has logged I_A , Π_A , sending of M_A , and M_B . Let dishonest U_B block all messages P_A sent to P_B but retains a copy of these incoming messages. P_B , having received no message from P_A within 2Δ time, will send Req_B for which TTP will respond by setting *outcome* = *aborted* and sending $Abort_{TTP}(B)$. Say U_B blocks $Abort_{TTP}(B)$ as well and crashes his node. The recovered P_A will find itself in having received only M_B and will send Res_A which, say, reaches TTP after *outcome* was set to *aborted*; TTP's response will be to send $Abort_{TTP}(A)$ to P_A . Let U_B re-boot N_B , delete Π_B from stable-store, and adjust the clock to make it appear as if the exchange phase is being executed for the first time. He replays the arrival of Π_B and of the blocked messages of P_A exactly at those instances when they arrived during the first execution. Since P_B 'receives' both M_A and $Ack_A(B)$, it delivers I_A to U_B .

Outline of Protocol P2

At the core of P1's inability to be crash tolerant is the fact that having both M_A and $Ack_A(B)$ is a sufficient condition for P_B to deliver I_A without consulting TTP at all. To remedy this, we need to add one more round to protocol P1 which P_X executes after having executed the first two rounds (concurrently) and check-pointed its state. In the third round, shown in Figure 4, P_A sends a second acknowledgement $Ack^2_A(B)$ to P_B indicating the reception of $Ack_B(A)$ and M_B .

Figure 4. The additional message round in P2.

If P_A does not receive $Ack_B^2(A)$ within 2Δ time after it has sent $Ack_A^2(B)$, P_A appeals to TTP by sending $First_Acks_A$ that contains both $Ack_A(B)$ and $Ack_B(A)$. TTP, after receiving $First_Acks_A$, sends $Ack_{TTP}^2(A)$ to P_A only if it has not earlier received Req_A (which would have resulted in sending $Abort_{TTP}(B)$ to P_B). Thus, the condition for normal termination of P_A is: *received*($Ack_{Z_A}^2(A)$), where $Z_A \in \{B, TTP\}$. By symmetry, the condition for normal termination of P_B is: *received*($Ack_{Z_B}^2(B)$), where $Z_B \in \{A, TTP\}$. Three messages used in addition to those used in P1 are shown in Table 2.

$Ack_A^2(B) = (L, A, H(N), H(Ack_B(A)), My_ack^2); A \rightarrow B$	$Ack_B^2(A) = (L, B, H(N), H(Ack_A(B)), My_ack^2); B \rightarrow A$
$First_Acks_A = (L, A, H(N), Ack_A(B), Ack_B(A)); A \rightarrow TTP$	$First_Acks_B = (L, B, H(N), Ack_B(A), Ack_A(B)); B \rightarrow TTP$
$Ack_{TTP}^2(A) = (L, TTP, H(N), H(Ack_A(B)), My_ack^2); TTP \rightarrow A$	$Ack_{TTP}^2(B) = (L, TTP, H(N), H(Ack_B(A)), My_ack^2); TTP \rightarrow B$

Table 2. Description of additional messages used in P2.

4.2. Malicious Abuser, Synchronous Communication (P2.1)

Protocol P2.1 is the same as P2, except that some of the messages used need to be changed:

1. $\Pi_A, M_A, M_{TTP}(A), \Pi_B, M_B$, and $M_{TTP}(B)$ of Table 1 change as mentioned in Section 3.2. $Ack_A(B)$ and $Ack_B(A)$ remain exactly as shown in Table 1.
2. In Table 2, $My_ack = K_B$ in $Ack_B^2(A)$ and in $Ack_{TTP}^2(A)$; similarly, $My_ack = K_A$ in $Ack_A^2(B)$ and $Ack_{TTP}^2(B)$.

Remark: *Simple Extension to Asynchronous Communication model.* P2 and P2.1 work correctly in the asynchronous model if TTP computes Δ during the start-up phase as described for P1 A and P1.1A.

4.3. Malicious and Restricted Abusers, Asynchronous Communication, on-line TTP (P2A and P2.1A)

These protocols are obtained by making P1 A and P1.1A crash-tolerant by having process P_X pessimistically checkpoint its state (as in P2 and P2.1) before processing a received message and by making TTP state-keeping. The

latter is done as follows: for any message received from P_X after $T_{EX} + 4\Delta$, TTP sends $Ack_Y(X)$ to P_X if both $Ack_A(B)$ and $Ack_B(A)$ have been received before $T_{EX} + 4\Delta$; otherwise it sends $Abort_{TTP}(X)$.

5. SUMMARY AND DISCUSSIONS

Table 3 summarises the protocols' characteristics using the following abbreviations:

- a dishonest user (DU) can be either a restricted abuser (RA) or a malicious abuser (MA),
- communication delay model (C) can be synchronous (S) or asynchronous (As),
- faults tolerated (FT) can be none (No) or crash (Cr), and
- the TTP can be either off-line (Off) or on-line (On), and either state-relinquishing (Sr) or state-keeping (Sk).

Protocol	DU	C	FT	TTP
P1	RA	S	No	Off/Sr
P1.1	MA	S	No	Off/Sr
P1A(P1.1A)	RA(MA)	As	No	On/Sr
P2	RA	S/As	Cr	Off/Sk
P2.1	MA	S/As	Cr	Off/Sk
P2A(P2.1A)	RA(MA)	As	Cr	On/Sk

Table 3. Characteristics of the Protocols Developed.

On-line TTP: state-relinquishing vs. state-keeping. Because of its simplicity, a non-fault-tolerant protocol with an on-line TTP is likely in practice, and a need may later arise to enhance the protocol to be crash-tolerant. Can an online-TTP protocol be made crash-tolerant simply through message logging, regardless of whether the delay model is synchronous or asynchronous? The answer depends on whether the non-fault-tolerant protocol keeps the TTP state-relinquishing or state-keeping. Our protocols P1A and P1.1A suggest that it is possible to design online-TTP protocols that keep the TTP state-relinquishing even with asynchronous delays. In making them crash-tolerant, we have to make the (on-line) TTP a state-keeping one (see P2A and P2.1A in Table 3). Thus, the answer to the above question is yes, provided the issue of state-relinquishing vs. state-keeping TTP is given due consideration.

Optimising the cost of message logging. The work by [4] proposes a logging-based approach to achieving crash-tolerant fair-exchange. It defines a point of no return for a user process, and claims that if the process would synchronously log all received messages before entering this point, then logging of other received messages can be done asynchronously. Their definition of point of no return suggests that there is only one such point for each process (Section 3.2 of [4]). We remark here that our protocol P2.1 without check-pointing and message logging is identical to the time-optimal protocol (scheme 3) of [5] which is asynchronous and non-fault-tolerant. Applying the approach of [4] to make this protocol crash-tolerant does *not* work: the point of no return for honest P_A turns out to be its sending M_A to P_B (see figure 2, round 1); suppose that N_A crashes in round 3 (see Fig. 4) after P_A has sent $Ack_A^2(B)$ but before $Ack_B(A)$ and M_B were to be logged by the asynchronous logging scheme; note that $Ack_A^2(B)$ contains K_A (see Section 4.2). Honest U_A suffers loss of fairness, even if U_B does not misbehave; to avoid loss of fairness, P_A must also log all received messages before sending $Ack_A^2(B)$. It appears that some protocols can have *multiple points* at which pessimistic logging is essential. Given that the number of messages received by a user process in a given execution is small, attempts to minimise the overhead of logging may not be worth the effort after all; further, every message received (and logged) may be useful later in any dispute resolution.

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