Assured-Timeliness Integrity Protocols for Distributable Real-Time Threads with in Dynamic Distributed Systems

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Abstract

Networked embedded systems present unique challenges for system designers composing distributed applications with dyanmic, real-time, and resilience requirements. We consider the problem of recovering from failures of distributable threads with assured timeliness in dynamic systems with overloads, and node and (permanent/transient) network failures. When a distributable thread encounters a failure that prevents its timely execution, the thread must be terminated. Thread termination involves detecting and aborting thread orphans, and delivering exceptions to the farthest, contiguous surviving thread segment for possible execution resumption. Thread termination operations must optimize system-wide timeliness. We present a scheduling algorithm called HUA and two thread integrity protocols called D-TPR and W-TPR. We show that they bound the orphan cleanup and recovery time with bounded loss of the best-effort property—i.e., high importance threads are always favored over low importance ones (for feasible completion), irrespective of thread urgency. Our implementation experience using the emerging Reference Implementation of Sun's Distributed Real-Time Specification for Java (DRTSJ) demonstrates the algorithm/protocols' effectiveness.

1. Introduction

In distributed systems, action and information timeliness is often end-to-end—e.g., a causally dependent, multi-node, sensor to shooter sequential flow of execution in network-centric warfare systems [3]. Designers and users of distributed systems often need to dependably reason about (specify, manage, predict) end-to-end timeliness. Many emerging such systems are being envisioned to be built using ad hoc network systems—e.g., those without a fixed network infrastructure and have dynamic node membership and network topology changes, including mobile, ad hoc wireless networks [2]. Reasoning about timeliness, especially end-to-end, is a very difficult and unsolved problem in such dynamic uncertain systems.

Maintaining end-to-end properties (e.g., timeliness, connectivity) of a control or information flow requires a model of the flow's locus in space and time that can be reasoned about. Such a model facilitates reasoning about, and resolving the contention for resources that occur along the flow's locus. The *distributable thread* abstraction which first appeared in the Alpha OS [13] and later in MK7.3 [17], OMG's Real-Time CORBA 1.2 [14], and Sun's emerging Distributed Real-Time Specification for Java (DRTSJ) [1] directly provides such a model as their first-class programming and scheduling abstraction. A distributable thread is a single thread of execution with a globally unique identity that extends and retracts through local and remote objects. We focus on distributable threads as our end-to-end control flow/scheduling abstraction, and hereafter, refer to them as *threads* except as necessary for clarity.

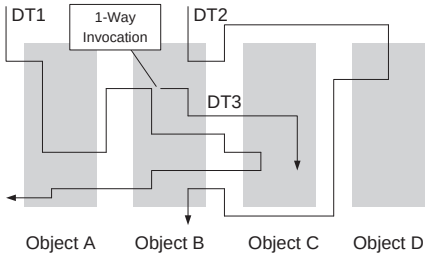


Figure 1. Distributable Threads

A thread carries its execution context as it transits node boundaries, including its scheduling parameters (e.g., time constraints, execution time), identity, and security credentials. The propagated thread context is intended to be used by node schedulers for resolving all node-local resource contention among threads such as that for node's physical

(e.g., CPU) and logical (e.g., locks) resources, according to a discipline that provides acceptably optimal system-wide timeliness. Figure 1 shows the execution of threads [14].

We consider the problem of developing assurances on thread timing behavior in dynamic systems, in the presence of application/network-induced uncertainties. The uncertainties include transient and sustained resource overloads (due to context-dependent thread execution times), arbitrary thread arrivals, node failures, and transient and permanent link failures (causing varying packet drop rate behaviors). Another distinguishing feature of motivating applications for this model (e.g., [3]) is their relatively long thread execution time magnitudes—e.g., milliseconds to minutes.

When overloads occur, meeting time constraints of all threads is impossible as the demand exceeds the supply. The urgency of a thread is sometimes orthogonal to the relative importance of the thread—-e.g., the most urgent thread may be the least important, and vice versa; the most urgent may be the most important, and vice versa. Hence when overloads occur, completing the most important threads irrespective of thread urgency is desirable. Thus, a distinction has to be made between urgency and importance during overloads. (During underloads, such a distinction generally need not be made—e.g., if all time constraints are deadlines, then EDF [8] can meet all deadlines.)

Deadlines cannot express both urgency and importance. Thus, we consider the *time/utility function* (or TUF) timeliness model [9] that specifies the utility of completing a thread as a function of its completion time. We specify a deadline as a binary-valued, downward "step" shaped TUF; Figure 2 shows examples. A thread's TUF decouples its importance and urgency—urgency is measured on the X-axis, and importance is denoted (by utility) on the Y-axis.

When thread time constraints are expressed with TUFs, the scheduling optimality criteria are based on maximizing accrued utility—e.g., maximizing the total thread accrued utility. Such criteria are called *utility accrual* (or UA) criteria, and sequencing (scheduling, dispatching) algorithms that optimize UA criteria are called UA sequencing algorithms (e.g., [5, 11]).

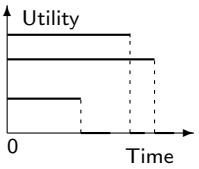


Figure 2. Step TUFs

UA algorithms that maximize total utility under downward step TUFs (e.g., [5, 11]) default to EDF during underloads, since EDF satisfies all deadlines during underloads. Consequently, they obtain the optimum total utility during underloads. During overloads, they inherently favor more important threads over less important ones (since more utility can be attained from the former), irrespective of thread urgency, and thus exhibit adaptive behavior and graceful timeliness degradation. This behavior of UA algorithms is called "best-effort" [11] in the sense that the algorithms strive their best to feasibly complete as many high importance threads — as specified by the application through TUFs — as possible.[1] Thus, high importance threads that arrive at any time always have a very high likelihood for feasible completion (irrespective of their urgency).

Our Contributions. When nodes transited by threads fail, it can divide those threads into several pieces. Segments of a thread that are disconnected from its node of origin (called the thread's root), are called *orphans*. When threads fail and cause orphans, application-supplied exception handlers must be released for execution on the orphan nodes. Such handlers may have time constraints themselves and will compete for their nodes' processor along with other threads. Under a termination model, when handlers execute (not necessarily when they are released), they will abort the associated orphans after performing recovery actions that are necessary to avoid inconsistencies. Once all handlers complete, thread execution can potentially be resumed from the farthest, contiguous surviving thread segment (from the thread's root). Such a coordinated set of recovery actions will preserve the abstraction of a continuous reliable thread.

Scheduling of the orphan-cleanup exception handlers along with threads must contribute to system-wide timeliness optimality. Untimely handler execution can degrade timeliness optimality—e.g.: high urgency handlers are delayed by low urgency non-failed threads, thereby delaying the resumption of high urgency failed threads; high urgency, non-failed threads are delayed by low urgency handlers.

A straightforward approach for scheduling handlers is to model them as traditional (single-node) threads, since these are similar in nature, with similar scheduling parameters such as execution time and time constraints. Further, the classical *admission control* strategy [6, 12, 16] can be used: When a thread T arrives on a node, if a feasible node schedule can be constructed such that it includes all the previously admitted threads and their handlers, besides T and its handler, then admit T and its handler; otherwise, reject. But this will cause the very fundamental problem that is solved by UA schedulers through their best-effort decision making—i.e., a newly arriving thread is rejected because it is infeasible, despite that thread being the most important. In contrast, UA schedulers will feasibly complete the high importance newly arriving thread (with high likelihood), at the expense of not completing some previously arrived ones, since they are now less important than the newly arrived.

In this paper, we consider the problem of recovering from thread failures with assured timeliness and best-effort property. We consider distributable threads that are subject to TUF time constraints. Threads may have arbitrary arrival behaviors, may exhibit unbounded execution time behaviors (causing node overloads), and may span nodes

[1] Note that the term "best effort" as used in the context of networks actually is intended to mean "least effort."

that are subject to arbitrary crash failures and a network with permanent/transient failures and unreliable transport mechanisms. For such a model, we consider the scheduling objective of maximizing the total thread accrued utility.

We present a UA scheduling algorithm called *Handler-assured Utility Accrual scheduling algorithm* (or HUA) for thread scheduling, and two protocols called *Decentralized Thread Polling with bounded Recovery* (or D-TPR) and *Wireless Thread Polling with bounded Recovery* (or W-TPR) for ensuring thread integrity. D-TPR targets networks with generally permanent network failures, and W-TPR targets mobile, ad hoc wireless networks with generally transient network failures. We show that HUA and D-TPR/W-TPR ensure that handlers of threads that encounter failures during their execution will complete within a bounded time, yielding bounded thread cleanup time. Yet, the algorithm/protocols retain the fundamental best-effort property of UA algorithms with bounded loss—i.e., a high importance thread that may arrive at any time has a very high likelihood for feasible completion. Our implementation experience using DRTSJ's emerging Reference Implementation (RI) demonstrates the algorithm/protocols' effectiveness.

Similar to UA algorithms, thread integrity protocols have been developed in the past—e.g., Thread Polling with bounded Recovery [6], Alpha's Thread Polling [13], Node Alive protocol [7]. However, none of these efforts provide time-bounded thread cleanup in the presence of node and (permanent/transient) network failures and unreliable transport mechanisms. Further, [6] suffers from unbounded loss of the best-effort property due to its admission control strategy (we show this in Section 3.3). In contrast, HUA and D-TPR/W-TPR provide bounded thread cleanup with bounded loss of the best-effort property in the presence of (permanent/transient) network failures and unreliable transport mechanisms — the first such algorithm/protocols. Thus, the paper's contribution is the HUA and D-TPR/W-TPR.

The rest of the paper is organized as follows: In Section 2, we state our models and objectives. Section 3 presents HUA, Section 4 presents D-TPR, and Section 5 presents W-TPR. In Section 6, we discuss our implementation experience. We conclude the paper in Section 7.

2. Models and Objectives

Threads. Threads execute in local and remote objects by location-independent invocations and returns. A thread begins its execution by invoking an object operation. The object and the operation are specified when the thread is created. The portion of a thread executing an object operation is called a *thread segment*. Thus, a thread can be viewed as being composed of a concatenation of thread segments.

A thread's initial segment is called its *root* and its most recent segment is called its *head*. The head of a thread is the only segment that is active. A thread can also be viewed as being composed of a sequence of *sections*, where a section is a maximal length sequence of contiguous thread segments on a node. A section's first segment results from an invocation from another node, and its last segment performs a remote invocation.

A section's execution time estimate is known when the thread arrives at the section's node. This execution time estimate includes that of the section's normal code and its exception handler code, and can be violated at run-time (e.g., due to context dependence, causing processor overloads).

A thread's total number of sections is unknown a-priori, as the thread is assumed to make remote invocations and returns based on context-dependent application logic.

The application is thus comprised of a set of threads, denoted $\mathbf{T} = \{T_1, T_2, T_3, \ldots\}$.

Timeliness Model. Each thread T_i's time constraint is specified using a TUF, denoted $U_i(t)$. A classical deadline is unit-valued—i.e., $U_i(t) = \{0, 1\}$, since importance is not considered. Downward step TUFs (Figure 2) generalize classical deadlines where $U_i(t) = \{0, \{n\}\}$. We focus on *non-increasing* (unimodal) TUFs, as they encompass the majority of time constraints of interest to us (e.g., [4]).

Each TUF U_i has an initial time I_i, which is the earliest time for which the function is defined, and a termination time X_i, which denotes the last point that the function crosses the X-axis. We assume that the initial time is the thread release time; thus a thread's absolute and relative termination times are the same. We also assume that $U_i(t) > 0, \forall t \in [I_i, X_i]$ and $U_i(t) = 0, \forall t \notin [I_i, X_i], \forall i$.

Abort Model. Each section of a thread has an associated exception handler. We consider a termination model for all thread failures. If a thread has not completed by its termination time, a time constraint-violation exception is raised, and handlers are released on all nodes hosting thread's sections. When a handler executes, it will abort the associated section after performing recovery actions that are necessary to avoid inconsistencies—e.g., rolling back/forward section's held logical and physical resources to safe states.

We consider a similar abort model for node and network failures. When a thread encounters a node/network failure causing orphans, an integrity protocol (e.g., D-TPR) delivers failure-exception notifications to all the orphan nodes. Those nodes then respond by releasing handlers which abort the orphans after executing recovery actions.

Each handler may have a time constraint, which is specified using a TUF. A handler's TUF's initial time is the time of failure of the handler's thread. The handler's TUF's termination time is relative to its initial time. Thus, a handler's absolute and relative termination times are *not* the same.

Each handler also has an execution time estimate. This estimate along with the handler's TUF are described by the handler's thread when the thread arrives at a node. Violation of the termination time of a handler's TUF will cause the immediate execution of system recovery code on that node, which will recover the thread section's held resources and return the system to a consistent and safe state.

System and Failure Models. We consider a system model, where a set of processing components, generically referred to as *nodes*, $N_i \in N, i \in [1, m]$, are interconnected via a network. We consider a multihop network model (e.g., WAN, MANET), with nodes interconnected through routers. Node clocks are synchronized—e.g., using [15].

Network is assumed to be unreliable. Nodes may fail arbitrarily by crashing (i.e., fail-stop). Network links may fail transiently or permanently, causing network partitions, again, arbitrarily. Only an unreliable message transport protocol like UDP is assumed. We describe thread integrity protocol-specific network assumptions in Sections 4 and 5.

Each node executes thread sections. The order of executing sections on a node is determined by the node scheduler. We consider Real-Time CORBA 1.2's [14] *Case 2* approach for thread scheduling. According to this approach, node schedulers use the propagated thread scheduling parameters and independently schedule thread sections to optimize the system-wide timeliness optimality criterion. Though this results in approximate, global, system-wide timeliness, Real-Time CORBA supports the approach due to its simplicity and capability for coherent end-to-end scheduling.

Scheduling Objectives. Our primary objective is to maximize the total thread accrued utility as much as possible. Further, the orphan cleanup and recovery time must be bounded. This is the time between the detection of a thread failure and the time at which all orphans of the thread complete. The algorithm must also exhibit the best-effort property of UA algorithms (Section 1) to the extent possible.

3. The HUA Algorithm

3.1. Rationale

Section Scheduling. Since the task model is dynamic—i.e., when threads will arrive at nodes, and how many sections a thread will have are statically unknown, node (section) schedules must be constructed solely exploiting the current system knowledge. Since the objective is to maximize the total thread accrued utility, a reasonable heuristic is a "greedy" strategy at each node: Favor "high return" thread sections over low return ones, and complete as many of them as possible before thread termination times, as early as possible (since TUFs are non-increasing).

The potential utility that can be accrued by executing a thread section on a node defines a measure of that section's "return on investment." We measure this using a metric called the *Potential Utility Density* (or PUD) [5]. On a node, a section's PUD measures the utility that can be accrued per unit time by immediately executing it on the node.

However, a section may encounter failures. We first define the concept of a *section failure* and a *released handler*:

Definition 1 (Section Failure)**.** *Consider a section S_i of a thread T_i. We say that S_i has failed when (a) S_i violates the termination time of T_i while executing, thereby raising a time constraint-violation exception on S_i's node; or (b) a failure-exception notification is received at S_i's node regarding the failure of a section of T_i that is upstream or downstream of S_i, which designates S_i as an "orphan-head."*

Definition 2 (Released Handler)**.** *A handler is released for execution when its section fails according to Definition 1.*

Since a section's best-case failure scenario is the absence of a failure for the section, the corresponding section PUD can be obtained as the utility accrued by executing the section divided by the time spent for executing the section. The section PUD for the worst-case failure scenario (one where the section fails, per Definition 1) can be obtained as the utility accrued by executing the handler of the section divided by the total time spent for executing the section and the handler.[2] The section's PUD can now be measured as the minimum of these two PUDs, as that is the worst-case.

Thus, on each node, HUA examines thread sections for potential inclusion in a feasible node schedule in the order of decreasing section PUDs. For each section, the algorithm examines whether that section and its handler can be feasibly completed (we discuss section and handler feasibility later in this subsection). If infeasible, the section and its handler are rejected. The process is repeated until all sections are examined, and the schedule's first section is dispatched for execution on the node.

[2] In the worst-case failure scenario, utility is accrued only for executing the section's handler; no utility is gained for executing the section, though execution time is spent for executing the section and its handler.

A section S_i that is rejected can be the head of its thread T_i; if so, S_i is reconsidered for scheduling on S_i's node, say N_i, until T_i's termination time expires.

If a rejected section S_i is not a head, then S_i's rejection is conceptually equivalent to the (crash) failure of N_i. This is because, S_i's thread T_i has made a downstream invocation after arriving at N_i and is yet to return from that invocation (that's why S_i is still a scheduling entity on N_i). If T_i had made a downstream invocation, then S_i had executed before, and hence was feasible and had a feasible handler at that time. S_i's rejection now invalidates that previous feasibility. Thus, S_i must be reported as failed and a thread break for T_i at N_i must be reported to have occurred to ensure system-wide consistency on thread feasibility. The algorithm does this by interacting with the integrity protocol (e.g., D-TPR).

This process ensures that the sections that are included in a node's schedule at any time have feasible handlers. Further, all their upstream sections also have feasible handlers on their respective nodes. Thus, when any such section fails (per Definition 1), its handler and the handlers of all its upstream sections will complete within a bounded time.

Note that no such assurances are afforded to sections that fail otherwise—i.e., the termination time expires for a section S_i, which has not completed its execution and is not executing when the expiration occurs. Thus, S_i and its handler are not part of the feasible schedule at the expiration time. For this case, S_i's handler is executed in a best-effort manner—i.e., in accordance with its potential contribution to the total utility (at the expiration time).

Feasibility. Feasibility of a section on a node can be tested by verifying whether the section can be completed on the node before the section's distributable thread's end-to-end termination time. Using a thread's end-to-end termination time for verifying the feasibility of a section of the thread may potentially overestimate the section's slack, especially if there are a significant number of sections that follow it in the thread. However, this is a reasonable choice, since the total number of sections of a thread is unknown. If a thread's total number of sections is known a-priori, then better schemes (e.g., [10]) that intelligently distribute the thread's total slack among all its sections can be considered.

For a section's handler, feasibility means whether it can complete before its *absolute* termination time, which is the time of thread failure plus the relative termination time of the section's handler. Since the thread failure time is impossible to predict, a reasonable choice for the handler's absolute termination time is the thread's end-to-end termination time plus the handler's termination time, as that will delay the handler's latest start time as much as possible. Delaying a handler's start time on a node is appropriate toward maximizing the total utility, as it potentially allows threads that may arrive later on the node but with an earlier termination time than that of the handler to be feasibly scheduled.

There is always the possibility that a new section S_i is released on a node after the failure of another section S_j at the node (per Definition 1) and before the completion of S_j's handler on the node. As per the best-effort philosophy, S_i must immediately be afforded the opportunity for feasible execution on the node, in accordance with its potential contribution to the total utility. However, it is possible that a schedule that includes S_i on the node may not include S_j's handler. Since S_j's handler cannot be rejected now, as that will violate the commitment previously made to S_j, the only option left is to not consider S_i for execution until S_j's handler completes, consequently degrading the algorithm's best-effort property. In Section 3.3, we quantify this loss.

3.2. Algorithm Overview

HUA's scheduling events at a node include the arrival of a thread at the node, release of a handler at the node, completion of a thread section or a section handler at the node, and the expiration of a TUF termination time at the node. To describe HUA, we define the following variables and auxiliary functions (at a node):

- $\mathcal{S}_r$ is the current set of unscheduled sections including a newly arrived section (if any). $S_i \in \mathcal{S}_r$ is a section. S_i^h denotes S_i's handler. T_i denotes the thread to which a section S_i and S_i^h belong. $S_i.X$ is S_i's termination time, which is the same as that of T_i's termination time. $S_i.ExT$ is S_i's estimated remaining execution time. $U_i(t)$ denotes S_i's TUF, which is the same as that of T_i's TUF. $U_i^h(t)$ denotes S_i^h's TUF.
- σ_r is the schedule (ordered list) constructed at the previous scheduling event. σ is the new schedule.
- H is the set of handlers that are released for execution on the node (per Definition 2), ordered by non-decreasing handler termination times. $H = \emptyset$ if all released handlers have completed.
- Function `updateReleaseHandlerSet()` inserts a handler S_i^h into H if the scheduler is invoked due to S_i^h's release; deletes a handler S_i^h from H if the scheduler is invoked due to S_i^h's completion. Insertion of S_i^h into H is at the position corresponding to S_i^h's termination time.
- `alertProtocol(`S_i`)` declares S_i as failed (e.g., with D-TPR, this is done by S_i's node not sending `POLL` messages to S_i's predecessor and successor section nodes).
- `IsHead(`S`)` returns true if S is a head; false otherwise. `headOf(`σ`)` returns the first section in σ.

- `sortByPUD(`σ`)` returns a schedule ordered by non-increasing section PUDs. If two or more sections have the same PUD, then the section(s) with the largest ExT will appear before any others with the same PUD.
- `Insert(`S`,`σ`,`I`)` inserts section S in the ordered list σ at the position indicated by index I; if entries in σ exists with the index I, S is inserted before them. After insertion, S's index in σ is I. `Remove(`S`,`σ`,`I`)` removes section S from ordered list σ at the position indicated by index I.
- `feasible(`σ`)` returns a boolean value indicating schedule σ's feasibility. σ is feasible, if the predicted completion time of each section S in σ, denoted $S.C$, does not exceed S's termination time. $S.C$ is the time at which the scheduler is invoked plus the sum of the ExT's of all sections that occur before S in σ and $S.ExT$.

```
1: input: S_r, σ_r, H; output: selected thread S_exe;
2: Initialization: t := t_cur; σ := ∅; HandlerIsMissed := false;
3: updateReleaseHandlerSet ();
4: for each section S_i ∈ S_r do
5:     if feasible(S_i)=false then
6:         reject(S_i);
7:     else S_i.PUD = min( U_i(t+S_i.ExT)/(S_i.ExT) , U_i^h(t+S_i.ExT+S_i^h.ExT)/(S_i.ExT+S_i^h.ExT) )
8: σ_tmp :=sortByPUD(S_r);
9: for each section S_i ∈ σ_tmp from head to tail do
10:    if S_i.PUD > 0 then
11:        Insert(S_i, σ, S_i.X);
12:        Insert(S_i^h, σ, S_i.X + S_i^h.X);
13:        if feasible(σ)=false then
               Remove(S_i, σ, S_i.X);
14:            Remove(S_i^h, σ, S_i.X + S_i^h.X);
15:            if IsHead(S_i)=false and S_i ∈ σ_r then
16:                alertProtocol(S_i);
17:    else break;
18: if H ≠ ∅ then
19:    for each section S^h ∈ H do
20:        if S^h ∉ σ then
21:            HandlerIsMissed := true;
22:            break;
23: if HandlerIsMissed := true then
24:    S_exe :=headOf(H);
    else
25:    σ_r := σ;
26:    T_exe:=headOf(σ);
27: return S_exe;
```

Algorithm 1: HUA: High Level Description

Algorithm 1 describes HUA at a high level of abstraction. When invoked at time t_{cur}, HUA first updates the set H (line 3) and checks the feasibility of the sections. If a section's earliest predicted completion time exceeds its termination time, it is rejected (line 6). Otherwise, HUA calculates the section's PUD (line 7).

The sections are then sorted by their PUDs (line 8). In each step of the *for*-loop from line 9 to line 17, the section with the largest PUD and its handler are inserted into σ, if it can produce a positive PUD. The schedule σ is maintained in the non-decreasing order of section termination times. Thus, a section S_i and S_i^h are inserted into σ at positions that correspond to $S_i.X$ and $S_i.X + S_i^h.X$, respectively.

If after inserting S_i and S_i^h into σ, σ becomes infeasible, S_i and S_i^h are removed (lines 13–14). If a section S_i that is removed is not a head and belonged to the schedule constructed at the previous scheduling event, the integrity protocol is notified regarding S_i's failure (lines 15–16).

If one or more handlers have been released but have not completed their execution (i.e., $H \neq \emptyset$; line 18), the algorithm checks whether any of those handlers are missing in the schedule σ (lines 19– 22). If any handler is missing, the handler at the head of H is selected for execution (line 24). If all handlers in H have been included in σ, the section at the head of σ is selected (line 26).

3.3. Algorithm Properties

Theorem 1. *If a section S_i fails (per Definition 1), then under HUA with zero overhead, its handler S_i^h will complete no later than $S_i.X + S_i^h.X$ (barring S_i^h's failure).*

Proof. If S_i violates the thread termination time at a time t while executing, then S_i is in HUA's current schedule. This implies that both S_i and S_i^h were feasible, and S_i^h was scheduled to complete by $S_i.X + S_i^h.X$.

If S_i receives a notification on the failure of an upstream section $\bar{S}_i$ at a time t, then all sections from $\bar{S}_i$ to S_i and their handlers are feasible on their nodes, as otherwise the thread execution would not have progressed to S_i. Thus, S_i^h is scheduled to complete by $S_i.X + S_i^h.X$. A similar argument holds for the case of S_i receiving a notification on the failure of a downstream section. Theorem follows. □

Consider a thread T_i that arrives at a node and releases a section S_i after the handler of a section S_j has been released on the node (per Definition 2) and before that handler (S_j^h) completes. Now, HUA *may* exclude S_i from a schedule until S_j^h completes, resulting in some loss of the best-effort property. To quantify this loss, we define the concept of a *Non Best-effort time Interval* (or NBI):

Definition 3. *Consider a scheduling algorithm $\mathcal{A}$. Let a section S_i arrive at a time t with the following properties: (a) S_i and its handler together with all sections in $\mathcal{A}$'s schedule at time t are not feasible at t, but S_i and its handler are feasible just by themselves; (b) One or more handlers (which were released before t) have not completed their execution at t; and (c) S_i has the highest PUD among all sections in $\mathcal{A}$'s schedule at time t. Now, $\mathcal{A}$'s NBI, $NBI_{\mathcal{A}}$, is defined as the duration of time that S_i will have to wait after t, before it is included in $\mathcal{A}$'s feasible schedule. Thus, S_i is assumed to be feasible together with its handler at $t + NBI_{\mathcal{A}}$.*

We now describe the NBI of HUA and other UA algorithms including DASA [5], LBESA [11], and AUA [6] (under zero overhead):

Theorem 2. *HUA's worst-case NBI is $t + \max_{\forall S_j \in \sigma_t} \left(S_j.X + S_j^h.X\right)$, where σ_t is HUA's schedule at time t. DASA's and LBESA's worst-case NBI is zero; AUA's is $+\infty$.*

Proof. The time t that will result in the worst-case NBI for HUA is when $\sigma_t = H \neq \emptyset$. Since S_i has the highest PUD and is feasible, S_i will be included in the feasible schedule σ, resulting in the rejection of some handlers in H. Consequently, the algorithm will discard σ and will select the first handler in H for execution. In the worst-case, this process repeats for each of the scheduling events that occur until all the handlers in σ_t complete. Since each handler in σ_t is scheduled to complete by $\max_{\forall S_j \in \sigma_t} \left(S_j.X + S_j^h.X\right)$, the earliest time that S_i becomes feasible is $t + \max_{\forall S_j \in \sigma_t} \left(S_j.X + S_j^h.X\right)$.

DASA and LBESA will examine S_i at t, since a task arrival is always a scheduling event for them. Further, since S_i has the highest PUD and is feasible, they will include S_i in their feasible schedules at t (before including any other tasks), yielding a zero worst-case NBI.

AUA will examine S_i at t, since a task arrival at any time is also a scheduling event under it. However, AUA is a TUF/UA algorithm in the classical admission control mould and will reject S_i in favor of previously admitted tasks, yielding a worst-case NBI of $+\infty$. □

Theorem 3. *The best-case NBI of HUA, DASA, and LBESA is zero; AUA's is $+\infty$.*

Proof. HUA's best-case NBI occurs when S_i arrives at t and the algorithm includes S_i and all handlers in H in the feasible schedule σ (thus the algorithm only rejects some *sections* in σ_t to construct σ). Thus, S_i is included in a feasible schedule at time t, resulting in zero best-case NBI.

The best-case NBI scenario for DASA, LBESA, and AUA is the same as their worst-case. □

4. The D-TPR Protocol

D-TPR targets systems with node and network failures that are generally permanent. The protocol is instantiated in a software component called the Thread Integrity Manager (or TIM). Every node hosting thread sections has a TIM, which continually runs D-TPR's polling operation.

The TIM's operations are considered to be administrative operations, and they are conducted with scheduling eligibility that exceeds all application threads. As a consequence, we ignore the (comparatively small, and bounded) processing delays on each node in the analysis. For simplicity in analysis, we also assume perfectly synchronized clocks.

Figure 3 shows a sequence diagram for the operation of D-TPR for a healthy thread.

4.1. Polling

Table 1. D-TPR Messages

Message	Contents	From/To
POLL	List of local section ID and remote section ID pairs. Remote section-IDs are either predecessor or successor sections to local section	travel back and forth between predecessor and successor nodes
NEW_HEAD	timed out section and predecessor section	node with upstream timeout to predecessor node
ENDORPHAN	timed out section and successor section	node with downstream timeout to successor node
ORPHANPROP	orphaned section and successor section	node with orphan section to successor node

At every polling interval t_p, the TIM on each node identifies the sections that are locally hosted. The TIM then sends a POLL message to each of its predecessor and successor nodes. Note that each node can host sections of several threads so a single node may have several predecessor and successor nodes.

Each POLL message (see Table 1) is a list of entries, where each entry contains a type, the local section ID the entry corresponds to, and a remote section ID. If the entry type is SUCCESSOR, the remote section ID will correspond to the successor section of the local section in the entry. Similarly, the remote section ID of PREDECESSOR corresponds to the predecessor section of the local segment in the entry. In this way, the node receiving the POLL message is able to discern (downstream or upstream) the message's origin and thus from which direction the section has been deemed healthy. This distinction becomes important for break detection and is discussed further.

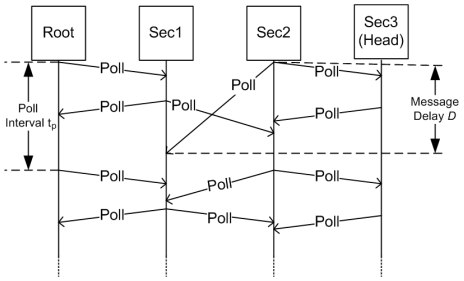


Figure 3. D-TPR Healthy Operation

4.2. Break Detection

When an invocation is made, D-TPR creates two timers which are set to a delay D. D is assumed to be the delay incurred in successfully transporting a message from one node to another in the network with very high probability, and is empirically determined (similar to our measurements in Section 6).[3] One timer is established for the downstream section and the other is established for the upstream section. The TIM on the node making the invocation (upstream side) creates a downstream-invocation timer that will cause a timeout when polling messages have not been received from downstream frequently enough. The TIM on the node hosting the remote object to which the invocation is being made (downstream side) creates an upstream-invocation timer that will cause a timeout when polling messages are not received from upstream frequently enough.

When a POLL message is received from upstream, the upstream-invocation timer is reset to D and resumes counting down. The same is true of the downstream-invocation timer when a POLL message is received from downstream. A "thread break" is declared when either the upstream or downstream-invocation time reaches zero. Recovery is different depending on which timer experiences the timeout.

Lemma 4. *Consider a section S_i and its successor section S_j. Under D-TPR, if S_j's node fails, or S_i becomes unreachable from S_j (but not necessarily vice versa), then S_i will detect a thread break between S_i and S_j within $t_p + D$.*

Proof. TPR's worst-case scenario for detecting this thread break occurs when S_j's node crashes immediately after S_j sends the POLL message to S_i (and the network successfully delivers that POLL to S_i), or when S_i becomes unreachable from S_j immediately after S_i receives S_j's POLL message. Consequently, S_i will miss discovering the thread break when it receives the POLL, and must wait for the lack of the next POLL from S_j to detect the break. The next POLL will be sent no later than one t_p, the lack of the receipt of which will be detected by S_i no later than one D. The lemma follows. □

Lemma 5. *Consider a section S_j and its predecessor S_i. Under D-TPR, if S_i's node fails, or S_j becomes unreachable from S_i (but not necessarily vice versa), then S_j will detect a thread break between S_i and S_j within $t_p + D$. S_j and its downstream sections are now said to be orphaned.*

Proof. The proof is similar to that of Lemma 4. □

[3]Thus, the lack of the receipt of a message at a destination node N_j within D of sending the message from a node N_i is considered a network failure—e.g., N_j is unreachable from N_i; a network partition between N_i and N_j—with high probability.

4.3. Recovery

D-TPR's recovery operations are administrative functions, and carries on below the level of application scheduling. While recovery proceeds, D-TPR activities continue concurrently. This allows the protocol to recognize and deal with multiple simultaneous breaks and cleanup operations.

If the upstream-invocation timer expires, the protocol assumes that the upstream section is unreachable and declares the local section associated with the timer to be an *orphan*. D-TPR then attempts to accomplish two things: first, force the upstream section to become the thread's *new head*; and second, force the downstream section to become an *orphan*.

To force the upstream section to become the *new head*, the protocol sends a NEW_HEAD message upstream and stops upstream POLL messages, which refresh the upstream section. If the upstream node receives the NEW_HEAD message, the upstream section will immediately begin behaving like a *new head*. If the upstream node does not receive the message, the upstream section's downstream-invocation timer will expire (due to the stopped POLL messages) forcing the section to become the *new head*.

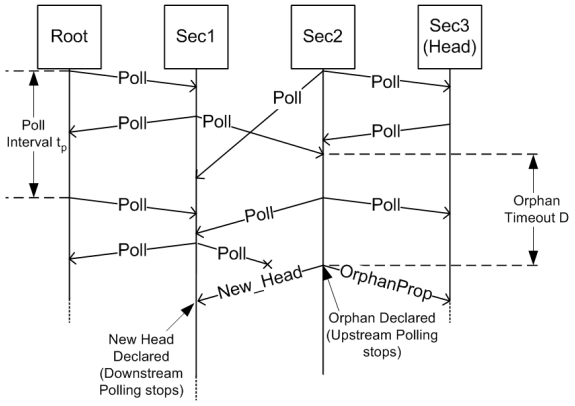


Figure 4. D-TPR Unhealthy Upstream

In order to force the downstream section to become an *orphan*, the protocol sends an ORPHANPROP message downstream and modifies its downstream POLL messages to include an orphan status. The downstream node will either receive the ORPHANPROP message and become an *orphan*, or the downstream section's timer will expire forcing it to become an *orphan*. When a section becomes an orphan, it propagates the ORPHANPROP message in order to identify all orphans. Figure 4 shows this scenario.

When a section's downstream-invocation timer expires, the protocol assumes that the downstream sections are unreachable and declares itself the *new head*of the thread. The *new head*then sends an ENDORPHAN downstream and ceases downstream refresh polling. In this way, the downstream section will either receive the ENDORPHAN notification and become an *orphan*or it's upstream timer will expire, making the section an orphan. Figure 5 shows this scenario.

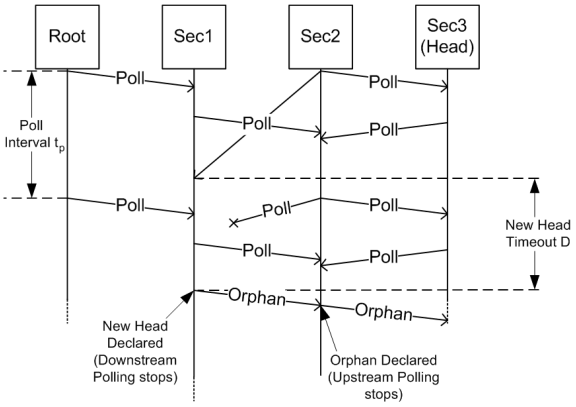


Figure 5. D-TPR Unhealthy Downstream

Lemma 6. *Under D-TPR, if a thread break occurs between S_i and its successor S_j, then S_i will become the new head within $t_p + 2D$. Since the new head of a thread is always directly upstream from a break, D-TPR therefore activates a new head within $t_p + 2D$.*

Proof. A thread break between S_i and S_j can occur in primarily two ways: (I) S_j's node fails, or S_i becomes unreachable from S_j; and (II) S_i's node fails, or S_j becomes unreachable from S_i. Lemma 4 identifies Case (I). Thus, S_i will detect the thread break within $t_p + D$, and immediately after, S_i will declare itself as the new head, within $t_p + D$. Case (II) is identified in Lemma 5. Thus, S_j will detect the break within $t_p + D$ and will send a NEW_HEAD message to S_i. Upon receipt of this message, S_i will declare itself as the new head, within a total of $t_p + 2D$ (after break detection), which is the worst-case. Lemma follows. □

Lemma 7. *Under D-TPR, if a thread break occurs between S_i and its successor S_j, then S_j will identify itself as an orphan within $t_p + 2D$.*

Proof. Proof is similar to that of Lemma 6. A thread break between S_i and S_j can occur in primarily two ways: (I) S_i's node fails, or S_j becomes unreachable from S_i; and (II) S_j's node fails, or S_i becomes unreachable from S_j. Case (I) is identified in Lemma 5. Thus, S_j will detect the break within $t_p + D$ and will immediately declare itself as an *orphan*, within $t_p + D$. Case (II) is identified in Lemma 4. Thus, S_i will detect the thread break within $t_p + D$, declare itself as the new head, and send an ENDORPHAN message to S_j. Upon receipt of this message, S_j will declare itself as an *orphan*, within a total of $t_p + 2D$ (after break detection), which is the worst-case. Lemma follows. □

4.4. Cleanup

A section that has been identified as an *orphan* will release the section's exception handler for aborting the section (i.e., the *orphan*) only if it has been designated an "orphan-head." This can happen in one of three ways: (1) The current head of the thread becomes an orphan; (2) A non-head orphan is returned to by an orphan-head and becomes a new orphan-head; and (3) An orphan's downstream-invocation timer expires forcing it to become a new orphan-head. .

Theorem 8. *Under D-TPR/HUA, if a thread break occurs between a section S_i and its successor S_j, then all orphans from S_j till the thread's current head S_{j+k}, for some $k \geq 1$, will be aborted in the LIFO-order—i.e., from S_{j+k} to S_j—and will complete by $t_p + (2+k)D + \sigma_{\alpha=0}^{k}(S_{j+\alpha}.X + S_{j+\alpha}^{h}.X)$, unless a section $S_{j+\alpha}$ becomes unreachable from $S_{j+\alpha+1}, 0 \leq \alpha \leq k-1$.*

Proof. Let the thread's execution sequence be: $\langle \cdots S_i, S_j, S_{j+1}, \cdots, S_{j+k} \rangle$. From Lemma 7, S_j will identify itself as an orphan within $t_p + 2D$. Following this, the ORPHANPROP message will be propagated from S_j to S_{j+k} within kD. Thus, S_{j+k} will become the first orphan-head and thus the first orphan to be aborted, followed by S_{j+k-1}, S_{j+k-2}, until S_j, following the LIFO-order, since $S_{j+k-\alpha}$ is always returned to by $S_{j+k-(\alpha-1)}, 0 \leq \alpha \leq k$ by the thread's execution sequence.

By Theorem 1, a section S_α's handler will complete within $S_\alpha.X + S_\alpha^h.X$, once it is an orphan-designate. Thus, all sections from S_j to S_{j+k} will complete within $t_p + 2D + kD + \sigma_{\alpha=0}^{k}(S_{j+\alpha}.X + S_{j+\alpha}^{h}.X)$. Theorem follows.

If a section $S_{j+\alpha}$ becomes unreachable from $S_{j+\alpha+1}$ $(0 \leq \alpha \leq k-1)$, then $S_{j+\alpha}$'s downstream invocation timer will expire before that of $S_{j+\alpha+1}$, designating $S_{j+\alpha}$ as an orphan-head before $S_{j+\alpha+1}$ — the theorem's exception. □

Theorem 9. *, Under D-TPR/HUA, if a thread breaks, then the thread's orphans will complete within a bounded time.*

Proof. This theorem follows from Theorem 8, except for the case when a section $S_{j+\alpha}$ becomes unreachable from $S_{j+\alpha+1}$ $(0 \leq \alpha \leq k-1)$ after a break occurs between S_i and its successor S_j. If $S_{j+\alpha}$ becomes unreachable from its successor $S_{j+\alpha+1}$, then $S_{j+\alpha}$'s downstream invocation timer will expire within $t_p + D$ (similar to Lemma 4, where $S_i \equiv S_{j+\alpha}$ and $S_j \equiv S_{j+\alpha+1}$), designating $S_{j+\alpha}$ as orphan-head. By Theorem 1, now $S_{j+\alpha}$ will cleanup within $t_p + D + S_{j+\alpha}.X + S_{j+\alpha}^{h}.X$. Theorem follows. □

5. The W-TPR Protocol

W-TPR is designed for mobile, ad hoc wireless networks, where communication is assumed to be unreliable and prone to transient failures (D-TPR considers communication failures to be permanent). The protocol exploits the fact that a thread is only adversely affected by a thread break if the head attempts to move across that break. In contrast, D-TPR detects a break and assumes that the break will be permanent; so it preempts the possibility of the head crossing the break by eliminating sections beyond the break point. W-TPR assumes that the breaks are not permanent.

W-TPR differs from D-TPR primarily in the way thread-breaks are determined. In D-TPR, breaks are recognized when communication between two consecutive nodes of a thread fails for longer than the message delay D, with very high probability. In W-TPR, breaks are never actually recognized. Instead, the protocol recognizes when communication errors affect either an invocation or a return (head movement) and provides maintenance accordingly.

Figure 6 shows the states and state transitions that a section undergoes in W-TPR. Note that no breaks are ever declared and that a section becomes an orphan only if it receives the ORPHAN message from an upstream section. Sections assume they are healthy until notified otherwise.

Downstream Head Movement. During an invocation, a thread section S_i makes a call on a remote object, which causes a second section, S_{i+1} to be created on the remote node. In order for the invocation to be successful, S_{i+1} must be created and S_i must be made aware of S_{i+1}.

When an invocation is made, an invocation request is sent downstream and the local section, S_i, begins waiting for invocation verification. The invocation is verified when the local section receives an INV-ACK from the downstream node or a POLL from the downstream node containing the section ID of

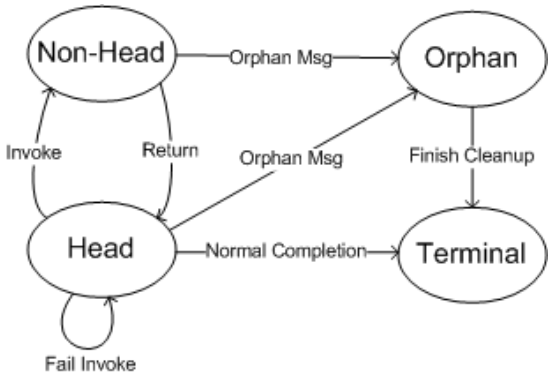


Figure 6. W-TPR Section State Diagram

the remote section (see further). When the invocation is verified, the local section is stopped until the remote section performs a return (head moves upstream).

Figure 7 shows an example of a successful invocation.

When the invocation is received by the downstream node, the downstream node attempts to finalize the invocation and sends an INV-ACK message to the upstream section (to quickly notify it of the successful receipt of the invocation message). Following this, the downstream node begins sending periodic POLL messages to the upstream section, at every polling interval t_p. When a healthy section receives a POLL message from an *orphan*, the healthy section returns an ORPHAN message to the *orphan*. If the *orphan* is not the *orphan-head*, similar to D-TPR, the ORPHAN message is propagated upstream.

Table 2 describes W-TPR messages.

The protocol resends the invocation request until either the invocation is verified, or the protocol deems that communication with the downstream node is not possible by waiting for an application-specified value t_n to expire and no INV-ACK or a POLL message is received from the downstream node during t_n. If communication with the downstream node is not possible, then the local section maintains head status and the application is notified that the invocation has failed. The TIM also sends an ORPHAN message downstream, in the event that a partial invocation was accomplished—i.e., the downstream node receives the invocation, and the upstream node becomes unreachable from the downstream node. Thus the downstream node's INV-ACK/POLL messages are not received upstream, while thread execution progresses on the downstream node and further downstream.

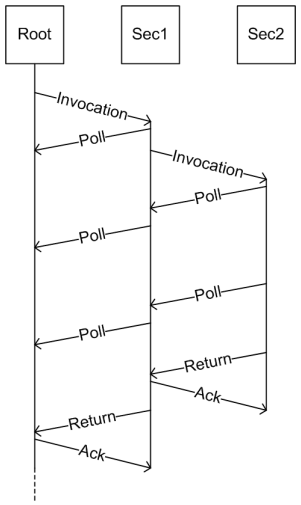


Figure 7. W-TPR Healthy Operation

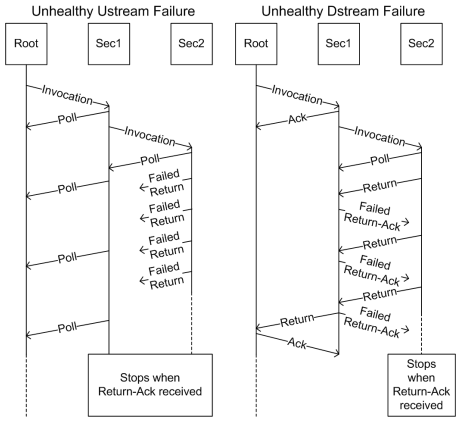


Figure 8. W-TPR Unhealthy Invocation

Figure 8 shows an unhealthy attempt at an invocation caused by an upstream failure on the left and a downstream failure on the right.

Lemma 10. *Under W-TPR, the location of a thread's head is ambiguous for at most t_n.*

Proof. Directly follows discussion. □

Upstream Head Movement. When the head is moving from the local node to an upstream node, the local node begins waiting for return verification from the upstream node. When the return message is received by the upstream node, the upstream node sends a return verification message RETURN-ACK downstream to the local node. If the verification is not received within t_n, then the return times-out (see Figure 9), and the protocol forces the return message to be resent. This process is repeated until the handshake successfully completes or the section on the local node violates its timing constraint. Even in the presence of upstream communication errors, the downstream section never becomes an *orphan*. Since the section has already finished executing and has a healthy return value, it would be fruitless to abort this section before delivering its return value.

Lemma 11. *Under W-TPR, a thread's head is never disconnected from the rest of the thread and no new head activation is required.*

Table 2. W-TPR Messages

Message	Contents	From/To
POLL	section ID pair of S_i and S_{i-1}	downstream node to upstream node
INV-ACK	section ID of section attempting invocation	downstream node to upstream node after receiving an invocation
RETURN-ACK	section ID of section attempting return	upstream node to downstream node after return success
ORPHAN	section ID pair of S_i (healthy) and S_{i+1} (orphan)	

Proof. Follows directly from the previous discussion. By Lemma 10, after t_n, the head moves downstream after a fully successful invocation. Any fully successful invocation can execute a return. If the upstream node becomes unreachable when the downstream node executes a return, the downstream section has completed its execution (hence it is returning) and is therefore not an *orphan*. □

Cleanup. A section becomes an *orphan* when it receives the ORPHAN message in response to one of its POLL messages. When the ORPHAN message is received, the section propagates that message downstream and waits for a return from its downstream section to be designated an orphan-head before starting cleanup, as in D-TPR. Cleanup begins when the furthest orphaned section is notified that it is an orphan.

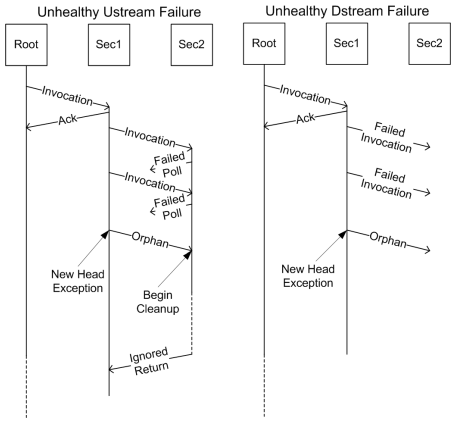


Figure 9. W-TPR Unhealthy Return

Theorem 12. *Under W-TPR, if a section S_i makes an unsuccessful invocation to its (potential) successor section S_j (i.e., S_j will be S_i's successor had if the invocation was successful), then all orphans that can potentially be created from S_j till the thread's furthest orphaned section $S_{j+k}, k \geq 1$, will be aborted in the LIFO-order and will complete within a bounded time under HUA, as long as no further failures occur between S_j and S_{j+k}.*

Proof. By Lemma 10, after t_n, S_i retains the head status since the invocation was unsuccessful, and an ORPHAN message is propagated to all downstream sections till S_{j+k}. The rest of the proof follows that of Theorem 9. □

Note that Theorem 12 holds only if no further failures occur between S_j and S_{j+k}. If such a failure were to occur, then the ORPHAN message may not be propagated or an orphan-head may not be able to return to a non-head orphan. D-TPR can detect such failures due to its continuous pairwise polling operation, whereas W-TPR is unable to do so precisely due its "on-demand" polling approach.

Theorem 13. *, Under W-TPR/HUA, if orphans are created for a thread as in Theorem 12, then all the orphans will complete within a bounded time, as long as no further failures occur between S_j and S_{j+k}.*

Proof. Follows from Theorem 12. □

6. Implementation Experience

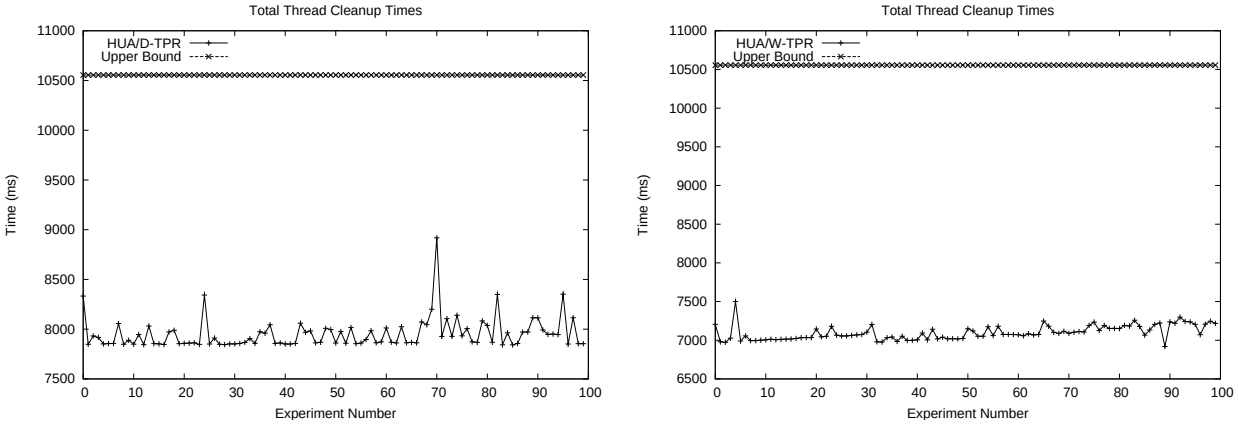
We implemented HUA, D-TPR, and W-TPR in DRTSJ's RI [1]. The RI includes a threads API, user-space scheduling framework for pluggable thread scheduling, and mechanisms for implementing thread integrity protocols (e.g., TIM). The RI infrastructure runs atop Apogee's Real-Time Specification for Java (RTSJ)-compliant Aphelion Java Virtual Machine. The RTSJ platform runs atop the Debian Linux OS (kernel version 2.6.16-2-686) on a 800MHz, Pentium-III machine. Our experimental testbed consisted of a network with five such DRTSJ nodes.

Our metrics of interest included Total Thread Cleanup Time and protocol overhead as measured by Thread Completion Time. We measured these during 100 experimental runs of our test application (application details are omitted for brevity). Each experimental run spawned a single distributable thread, which propagated to five other nodes and then returned back through the same five nodes.

The Total Thread Cleanup Time is the time between the failure of a thread's node or communication link and the completion of the handlers of all the orphan sections of the thread. Figures 10(a) and 10(b) show the measured cleanup times for HUA/D-TPR and HUA/W-TPR, respectively. The cleanup times are plotted against the protocols' cleanup upper bound times for the thread set used in our experiments. From the figures, we observe that both HUA/D-TPR and HUA/W-TPR satisfy their cleanup upper bound, thereby validating Theorems 9 and 13.

Thread Completion Time is the difference between when a root section starts execution and when it completes. Thread cleanup time is not taken into consideration here. Figures 11(a) and 11 show the thread completion times of experiments 1) with failures and D-TPR/W-TPR, 2) without failures but with D-TPR/W-TPR, 3) without failures and without D-TPR/W-TPR, and 4) with failures but without D-TPR/W-TPR. By measuring the thread completion times under these scenarios, we measure the overhead each protocol incurs in terms of the increase in thread completion times.

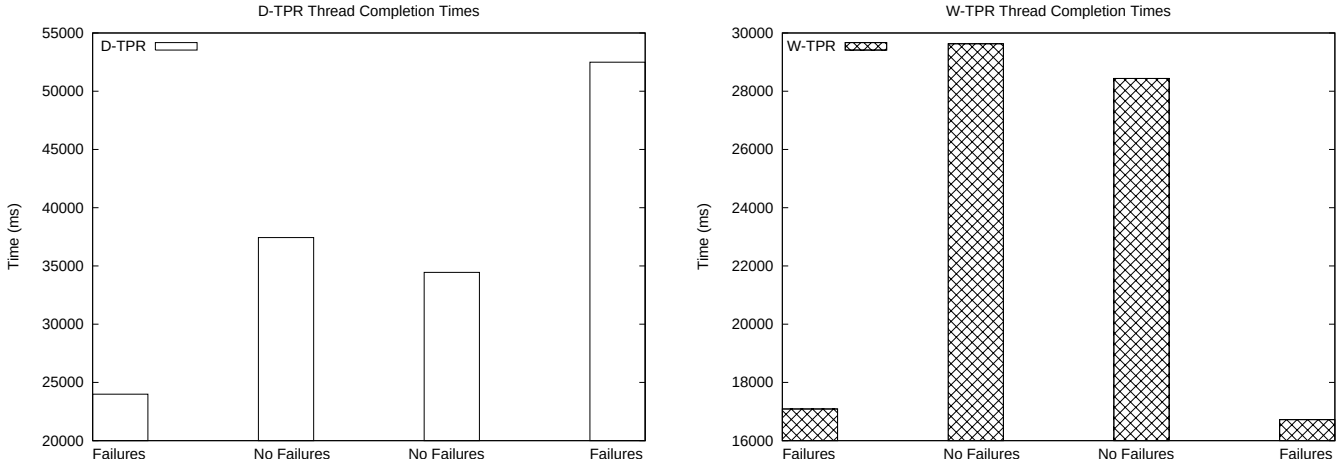
Figure 11(a) shows the completion times for experiments with and without D-TPR. We observe that the completion times of successful threads without D-TPR is smaller than that with D-TPR. This is to be expected as D-TPR incurs a



(a) D-TPR Thread Cleanup Times (b) W-TPR Thread Cleanup Times

Figure 10. Thread Cleanup Times for D-TPR and W-TPR

non-zero overhead. However, we also observe that the completion times of failed threads with D-TPR are shorter than even the completion times of successful threads without D-TPR. This is because, orphan cleanup can occur in parallel with the continuation of a repaired thread, allowing the repaired thread to finish without waiting for all orphans to run to completion. A successful thread, on the other hand, must wait for all sections to finish before it can complete, increasing its completion time. Figure 11(a) also shows that failed threads with D-TPR complete much more quickly than failed threads with no D-TPR support.



(a) D-TPR Thread Completion Times (b) W-TPR Thread Completion Times

Figure 11. W-TPR Thread Completion Times

Figure 11 shows completion times for experiments run with and without W-TPR. As the figure shows, the measurements taken in the absence of W-TPR are only slightly lower than the measurements taken in the presence of W-TPR. We observe that W-TPR incurs relatively little overhead while providing the properties discussed in Section 5.

7. Conclusions and Future Work

We present a real-time scheduling algorithm called HUA and two protocols called D-TPR and W-TPR. The algorithm/protocols' consider distributable threads and their exception handlers with TUF time constraints. We show that HUA and D-TPR/W-TPR bound the orphan cleanup and recovery time with bounded loss of the best-effort property — the first such algorithm/protocols for systems with (permanent/transient) network failures and unreliable transport. Our implementation using the emerging DRTSJ/RI demonstrates the algorithm/protocols' effectiveness.

Directions for future work include allowing threads to share non-CPU resources, establishing assurances on thread time constraint satisfactions', and extending results to arbitrary graph-shaped, multi-node, causal control/data flows.

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