A method for verifying liveness of protocols modeled as a class of ECFSM

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Abstract

In a previous work, we proposed a method for verifying safety of communication protocols modeled as two extended communicating finite-state machines with two unbounded FIFO channels connecting them. This paper presents a method for verifying liveness based on the above method. Liveness is formulated as Q -liveness which states “ $\forall gs \in RS \exists gs' \in GS_Q \{gs' \text{ is reachable from } gs\}$ ”, where RS and GS_Q denote the set of reachable global states and the set of global states satisfying the property Q , respectively. In the proposed verification method, a finite degenerated reachability graph, DRG , of a given protocol is constructed. In DRG , each node represents a subset of reachable states, and if there exists an edge from a node v_i to another node v_j , where v_i and v_j represent subsets of reachable states RS_i and RS_j respectively, then “ $\forall gs \in RS_i \exists gs' \in RS_j \{gs' \text{ is reachable from } gs\}$ ” holds. By exploring DRG , Q -liveness is shown to hold. An experimental result on verifying liveness of a sample protocol extracted from the data transfer phase of the OSI session protocol, is also described to show the effectiveness of the verification method.

Keyword Codes: C.2.2;D.2.4

Keywords: Network Protocol; Program Verification

1. Introduction

For communication protocols, safety and liveness are considered as fundamental properties, and methods for verifying those properties have been studied to obtain reliable communication systems^[1]. In general, the safety property states that the protocol never falls into any undesirable state, and the liveness property states that some desirable event will eventually occur. As an instance of desirable events, we can consider “receptions of some specified message” or “transitions to normal termination state”.

2-CFSMs (resp. 2-ECFSMs) is a protocol model consisting of two communicating finite-state machines (resp. two extended communicating finite-state machines) and bilateral FIFO queues connecting them. A state of an extended communicating finite-state machine is expressed by not only a value of finite control but also context variables whose

ranges of values are potentially infinite. If the boundedness of the communication channels of 2-CFSMs is guaranteed, the set of global states is finite and therefore many important properties, including safety and liveness, are decidable in principle^[2]. Some decision procedures have been proposed^{[3][4]} for channel-bounded 2-CFSMs based on state exploration which enumerates reachable global states. However, there exist many practical protocols whose state spaces are infinite because of the unboundedness of communication channels or infinite state space of individual protocol machines. Since both non-channel-bounded 2-CFSMs and a two counter machine, which is a subclass of single ECFSM, are known to be universal^{[2][5]}, traditional state exploration techniques cannot be used for such a class. For non-channel-bounded CFSMs, several verification methods were proposed^{[6][7][8]}. Among them, only [8] discussed a method for verifying liveness.

On the other hand, as a method for verifying safety of non-channel-bounded 2-ECFSMs, we proposed a verification method via invariant^{[9][10]} under the assumption that context variables are restricted to integral registers and the operations on integral registers are restricted to substitution, addition and comparison of two values. In spite of the assumption, a wide variety of practical protocols which use “sequence numbers”, can be modeled as 2-ECFSMs. On the other hand, since a class of two-counter machines is still a subclass of such restricted ECFSM, the expressive power of such restricted 2-ECFSMs is much higher than 2-CFSMs. The verification method proposed in ^{[9][10]} is summarized as follows:

- (a) A human verifier describes a logical formula, say F , on global states which is expected to satisfy (i) $RS \subseteq GS(F)$ and (ii) $GS(F) \subseteq SAFE$, where RS , $GS(F)$ and $SAFE$ denote the reachable global states, the set of global states satisfying F and the set of safe global states, respectively.
- (b) A verification system automatically shows the above (i) and (ii) by structural induction on event sequences.

In this paper, a method for verifying liveness of 2-ECFSMs is proposed by extending the above method for verifying safety. Liveness property is formulated as Q -liveness which states “ $\forall gs \in RS \exists gs' \in GS_Q \{gs' \text{ is reachable from } gs\}$ ”, where RS and GS_Q denote the set of reachable states and the set of global states satisfying the property Q , respectively. In the proposing verification method of liveness, a finite degenerated reachability graph DRG of a given protocol is constructed. In DRG , each node v_i represents a subset RS_i of reachable states. For each RS_i , all states in RS_i share common properties concerned with the followings:

- the states of the finite controls of the protocol machines,
- characteristics of message type sequences in the channels (e.g. “ $A \cdot A^*$ ”, one or more type A messages),
- characteristics of parameters of messages in the channels (e.g. they are arranged in strictly increasing order), and
- relations between the values of the context variables of the protocol machines and parameters of messages in the channels.

The verification method for safety in [9] gives a way to define such subsets. In DRG , if there exists an edge from a node v_i to another node v_j , then “ $\forall gs \in RS_i \exists gs' \in RS_j \{gs' \text{ is reachable from } gs\}$ ” holds. By exploring DRG , Q -liveness is shown to hold.

The rest of this paper is organized as follows: Section 2 defines a class of 2-ECFSMs. Section 3 summarizes a verification method of safety, proposed in [9]. In section 4, we define Q -liveness and degenerated reachability graphs, and give a method for constructing a degenerated reachability graph DRG . Also in section 4, we discuss another restriction on protocol model to obtain an efficient verification procedure. In section 5, a verification system and an experimental result on verifying liveness of a sample protocol are described.

2. Protocol Model

Two extended communicating finite-state machines (2-ECFSMs) are a protocol model which consists of two extended communicating finite-state machines and two unbounded FIFO channels connecting them. Formally, it is defined as follows.

Definition 1 A *protocol machine* PM is a 4-tuple (S, Σ, T, SI) , where

- (M1) $S = \langle SF, r \rangle$ defines a set of *states*, where SF is the state set of the finite control part of the machine and r is the number of registers (context variables) which store nonnegative integers. Let $\mathcal{N}$ denote the set of nonnegative integers. The state space of the protocol machine is $SF \times \mathcal{N}^r$.
- (M2) $\Sigma = \Sigma_- \cup \Sigma_+$: a finite set of *message types*. Σ_- is a set of message types which PM can send and Σ_+ is a set of message types which PM can receive. Σ_- and Σ_+ are supposed to be disjoint. For $d \in \Sigma$ and $n \in \mathcal{N}$, $\langle d, n \rangle$ is called a *message* and n is called the *parameter* of the message. The number of parameters of a message is assumed to be exactly one only for simplicity. In the following, for a message sequence u , $type(u)$ and $parameter(u)$ denote the message type sequence and the parameter sequence of u , respectively. The set of *events* EV of the protocol machine is defined in connection with the set of messages sent or received by the machine, i.e. $EV = \{\langle d, n \rangle \mid d \in \Sigma_-, n \in \mathcal{N}\} \cup \{\langle d, n \rangle \mid d \in \Sigma_+, n \in \mathcal{N}\}$. The former subset is the set of *sending events* and the latter one is the set of *receiving events*. To explicitly express that the event $\langle d, n \rangle$ is a sending (or receiving) event, we may write $-\langle d, n \rangle$ (or $+\langle d, n \rangle$).
- (M3) T : a set of *actions*. An action (sf, d, sf', C, R) is a 5-tuple defining a set of transitions, where $sf \in SF$, $d \in \Sigma$, $sf' \in SF$. C , called an *action condition*, is a simultaneous inequality to be satisfied by the parameter value n of the outgoing or incoming message and the register values $p_1, p_2, \dots, p_r$ of the protocol machine at the state just before the action. R , called a *register modification function* is a partial function from $\mathcal{N}^r \times \mathcal{N}$ to $\mathcal{N}^r$ defining the changes of register values in the transition. For a state $s = (sfc, p_1, p_2, \dots, p_r)$ of the protocol machine and an action $(sf, d, sf', C, R) \in T$, if $sfc = sf$ and $p_1, p_2, \dots, p_r, n$ satisfy C , then the protocol machine can send or receive (according to d in Σ_- or Σ_+) the message $\langle d, n \rangle$ at the state s and then the machine transits to $s' = (sf', R(p_1, p_2, \dots, p_r, n))$. A nondeterministic *state transition function* δ from $SF \times \mathcal{N}^r \times EV$ to $SF \times \mathcal{N}^r$ is determined by T as follows, where $s = (sf, p_1, p_2, \dots, p_r)$ and $e = \langle d, n \rangle$:

$$\begin{aligned} \delta(s, e) = \{ & (sf', R(p_1, p_2, \dots, p_r, n)) \mid \\ & (sf, d, sf', C, R) \in T \text{ and } p_1, p_2, \dots, p_r, n \text{ satisfies } C \}, \end{aligned}$$

For $s \in SF \times \mathcal{N}^r$ and $e \in EV$, if $\delta(s, e)$ is not empty, then the event e is said to be *executable* in the state s .

(M4) $SI \subseteq SF \times \mathcal{N}^r$: a set of *initial states*. $\square$

Definition 2 For two protocol machines $PM_A = (\langle SF_A, \tau_A \rangle, \Sigma_A, T_A, SI_A)$ and $PM_B = (\langle SF_B, \tau_B \rangle, \Sigma_B, T_B, SI_B)$, if $\Sigma_{B-} = \Sigma_{A+}$ (denoted by Σ_{BA}) and $\Sigma_{A-} = \Sigma_{B+}$ (denoted by Σ_{AB}), $\Pi = (PM_A, PM_B)$ is called a *protocol*. A 4-tuple $\langle s_A, s_B, u_{BA}, u_{AB} \rangle \in \langle SF_A \times \mathcal{N}^{r_A}, SF_B \times \mathcal{N}^{r_B}, \langle \Sigma_{BA} \times \mathcal{N} \rangle^*, \langle \Sigma_{AB} \times \mathcal{N} \rangle^* \rangle$ is called a *global state* of the protocol Π . u_{BA} and u_{AB} denote message sequences in the channels from PM_B to PM_A and from PM_A to PM_B respectively. A global state $\langle si_A, si_B, \varepsilon, \varepsilon \rangle$ (ε is the empty sequence) is called an *initial global state* of Π , when $si_A \in SI_A$, and $si_B \in SI_B$, or when $\langle si_A, si_B \rangle$ is in a specified subset of $SI_A \times SI_B$ if such a subset is explicitly specified. $\square$

Definition 3 A global state $gs' = \langle s'_A, s'_B, u'_{BA}, u'_{AB} \rangle$ is said to be *directly reachable* from $gs = \langle s_A, s_B, u_{BA}, u_{AB} \rangle$ (denoted by $gs \rightarrow gs'$) if one of the following conditions is satisfied for some $d \in \Sigma_{BA} \cup \Sigma_{AB}$ and $n \in \mathcal{N}$.

$$\begin{aligned} s'_A &\in \delta_A(s_A, -\langle d, n \rangle), \quad s'_B = s_B, \quad u'_{BA} = u_{BA}, \quad \text{and} \quad u'_{AB} = u_{AB} \cdot \langle d, n \rangle; \\ s'_A &\in \delta_A(s_A, +\langle d, n \rangle), \quad s'_B = s_B, \quad \langle d, n \rangle \cdot u'_{BA} = u_{BA}, \quad \text{and} \quad u'_{AB} = u_{AB}; \\ s'_A &= s_A, \quad s'_B \in \delta_B(s_B, -\langle d, n \rangle), \quad u'_{BA} = u_{BA} \cdot \langle d, n \rangle, \quad \text{and} \quad u'_{AB} = u_{AB}; \\ s'_A &= s_A, \quad s'_B \in \delta_B(s_B, +\langle d, n \rangle), \quad u'_{BA} = u_{BA}, \quad \text{and} \quad \langle d, n \rangle \cdot u'_{AB} = u_{AB}. \end{aligned}$$

If one of the former (or the latter) two conditions holds, then PM_A (or PM_B) is called the *performer*, the relation is also denoted $gs \xrightarrow{A, \langle d, n \rangle} gs'$ (or $gs \xrightarrow{B, \langle d, n \rangle} gs'$), and $(\langle d, n \rangle, A)$ (or $(\langle d, n \rangle, B)$) is said to be *executable* in gs . The reflexive transitive closure of the relation “ $\rightarrow$ ” is denoted by “ $\xrightarrow{*}$ ”. If $gs \xrightarrow{*} gs'$, then gs' is said to be *reachable* from gs . $\square$

For a protocol $\Pi = (PM_A, PM_B)$, the set of all reachable global states from some initial global state is called the *reachability set* of Π .

3. A Method for Verifying Safety

The verification method for liveness proposed in this paper is based on the method for verifying safety proposed in [9]. Before describing the method for verifying liveness, we will review the method for verifying safety.

For a protocol Π , if a logical formula F holds for every reachable global state, F is said to be an invariant in Π . Let $GS(P)$ denote the set of global states satisfying a formula P . In the verification method, the verification system shows that (a) a logical formula F , described by a human verifier, is an invariant in a protocol Π , and (b) $GS(F)$ does not contain any unsafe state. If (a) and (b) hold, we can conclude that Π is safe.

3.1 Safety Property

If the reachability set of a protocol Π does not contain the following states, Π is said to be *safe*:

- A global state $gs = \langle s_A, s_B, u_{BA}, u_{AB} \rangle$ is called a *deadlock state* if $u_{BA} = u_{AB} = \varepsilon$ and any sending event is executable in neither s_A nor s_B ;

- A global state $gs = \langle s_A, s_B, u_{BA}, u_{AB} \rangle$ is called an *unspecified reception state* if either $u_{BA} \neq \varepsilon$ and $\delta_A(s_A, +head(u_{BA}))$ is undefined, or $u_{AB} \neq \varepsilon$ and $\delta_B(s_B, +head(u_{AB}))$ is undefined, where for a nonempty sequence α , $head(\alpha)$ denotes the first element of α .

A deadlock state and an unspecified reception state are called *unsafe states*.

3.2 Description of a Logical Formula

In what follows, assume that a given formula F is in a disjunctive normal form $F = P_1 \vee P_2 \vee \dots \vee P_n$. Each disjunct P_i of the formula F is a conjunction of *atomic formulae* (or simply *atoms*) of the following four types. An example of P_i is found in Table 4.

- (AF1) A pair of finite states $\langle sf_A, sf_B \rangle$, where $sf_A \in SF_A$ and $sf_B \in SF_B$, is an atom which holds for a global state $\langle s_A, s_B, u_{BA}, u_{AB} \rangle$ iff the finite control parts of s_A and s_B are sf_A and sf_B respectively. We assume that exactly one AF1 type atom appears in each disjunct P_i .
- (AF2) A pair of predicates $\langle P_{BA}, P_{AB} \rangle$ to specify sets of message type sequences in two communication channels is an atom which holds for a global state $\langle s_A, s_B, u_{BA}, u_{AB} \rangle$ iff P_{BA} holds for u_{BA} and P_{AB} holds for u_{AB} . Such a predicate may be defined by a regular expression. For instance, an AF2 type atom $\langle \varepsilon, MIP \cdot MIP^* \rangle$ states that $type(u_{BA}) \in L(\varepsilon)$ and $type(u_{AB}) \in L(MIP \cdot MIP^*)$, where $L(r)$ denotes the set of sequences expressed by the regular expression r . We also assume that exactly one AF2 type atom appears in each P_i .
- (AF3) For a message sequence u and a set of message types Σ_X , $subseq(u, \Sigma_X)$ denotes the subsequence of u whose message types are restricted to Σ_X . A predicate on a message parameter sequence of $subseq(u_{BA}, \Sigma_X)$ (or $subseq(u_{AB}, \Sigma_X)$) for some Σ_X is also an atom. For instance, “ $step1(subseq(u_{AB}, \{MIP\}))$ ” states that the parameter sequence $parameter(subseq(u_{AB}, \{MIP\}))$ satisfies the predicate “ $step1$ ”. By “ $step1$ ” we mean that the parameter sequence is an increasing sequence such that the difference between every adjacent elements is one.

Predicates which appear in AF2 or AF3 type atoms are defined in terms of conditional rewriting rules and conditional inequalities.

- (AF4) A linear inequality which represents the relation on the values of registers of protocol machines and the parameter values of messages in communication channels. For instance, “ $VM_A = lastp(subseq(u_{AB}, \{MIP\})) + 1$ ” is an AF4 type atom which states that the value of the register VM of PM_A is equal to the parameter of the last MIP message in u_{AB} plus 1 at the global state under consideration.

For the disjunct P_i , let $sf_A(P_i)$ and $sf_B(P_i)$ denote the states of finite control of PM_A and PM_B respectively, specified in the AF1 type atom in P_i .

3.3 An Example

For example, let us consider a protocol $\Pi_{EX} = (PM_A, PM_B)$ based on the data transfer phase of OSI session protocol, and a formula F_{EX} (Table 1, 2.)

Table 1: Registers of the protocol machines

register name	meaning
VM_A, VM_B	the next serial number to be used
VA_A, VA_B	the lowest serial number to which a synchronization point confirmation is expected

Table 2: The definition of the actions of the protocol machines

$ta1 = (STA713_A, \langle MIP, sn \rangle, STA713_A, C_{ta1}, R_{ta1}) \in T_A$ $C_{ta1} : \{sn = VM_A\}$ $R_{ta1}(VM_A, VA_A) = (VM_A + 1, VA_A)$
$ta2 = (STA713_A, \langle MIA, sn \rangle, STA713_A, \phi, R_{ta2}) \in T_A$ $R_{ta2}(VM_A, VA_A) = (VM_A, sn + 1)$
$tb1 = (STA713_B, \langle MIA, sn \rangle, STA713_B, C_{tb1}, R_{tb1}) \in T_B$ $C_{tb1} : \{VM_B \geq VA_B + 1, VM_B = sn + 1, sn \geq VA_B\}$ $R_{tb1}(VM_B, VA_B) = (VM_B, sn + 1)$
$tb2 = (STA713_B, \langle MIP, sn \rangle, STA713_B, \phi, R_{tb2}) \in T_B$ $R_{tb2}(VM_B, VA_B) = (VM_B + 1, VA_B)$

A protocol machine PM_A can set up a synchronization point with a serial number sn in data stream by sending a message $\langle MIP, sn \rangle$ to the peer protocol machine PM_B . PM_A increments a serial number of synchronization point by one on every transmission of message $\langle MIP, sn \rangle$. The other protocol machine PM_B can confirm to synchronization point by sending a message $\langle MIA, sn \rangle$. PM_B does not need to respond every serial number. The serial number sn of message $\langle MIA, sn \rangle$ is equal to the maximum serial number among messages $\langle MIP, x \rangle$ which PM_B have already received.

Protocol machines PM_A and PM_B are defined as follows:

- The finite controls of PM_A and PM_B consist of one state $STA713_A$ and $STA713_B$ respectively. Both PM_A and PM_B have two integral registers. Table 1 tells intuitive description of those registers.
- $\Sigma_{AB} = \{MIP\}$, $\Sigma_{BA} = \{MIA\}$.
- Table 2 defines the set of actions of PM_A and PM_B . For example, t_3 states the followings:
 - (i) If the two integral registers VM_B and VA_B satisfy the relation $VM_B > VA_B$, then PM_B can send a message $\langle MIA, sn \rangle$ such that $VM_B = sn + 1$, $sn \geq VA_B$.
 - (ii) By executing the action the integral register VA_B is incremented by one.
- The initial states of PM_A and PM_B are $\langle sf_A, 0, 0 \rangle$ and $\langle sf_B, 0, 0 \rangle$ respectively.

A human verifier is able to divide the reachability set of Π_{EX} into 8 distinguished subsets by considering the truth values of the following statements: (1) There exists some messages whose message types are MIP, on the communication channel from PM_A to PM_B ; (2) There exist some messages whose message types are MIA, on the communication channel from PM_B to PM_A ; (3) PM_B can send a message whose message type is MIA, that is equivalent to the inequality $VM_B > VA_B$.

Table 3: Predicates and functions for a message sequence

name	meaning
$step1(\alpha)$	α is an increasing sequence by step 1
$st-inc(\alpha)$	α is a strictly increasing sequence
$firstp(\alpha)$	the parameter value of the first message of α
$lastp(\alpha)$	the parameter value of the last message of α

We define some predicates and functions to describe a formula. Table 3 shows an example of a definition of them. Table 4 shows an example of a logical formula that characterizes the above-mentioned eight subsets of the global states of Π_{EX} .

3.4 A Verification Procedure

In the verification method, for a given protocol Π and a formula $F = P_1 \vee P_2 \vee \dots \vee P_n$, F is shown to be an invariant of Π by the structural induction on event sequences of Π as follows.

Inductive basis: Prove that every initial global state of Π satisfies F .

Inductive step: Prove that

$$\forall gs \in GS(F) \forall gs'(gs \rightarrow gs') \{gs' \in GS(F)\}. \quad (*1)$$

Observe that $GS(F) = GS(P_1) \cup GS(P_2) \cup \dots \cup GS(P_n)$. Therefore, (*1) is equivalent to

$$\forall i(1 \leq i \leq n) \forall gs \in GS(P_i) \forall gs'(gs \rightarrow gs') \exists j(1 \leq j \leq n) \{gs' \in GS(P_j)\}.$$

Thus (*1) is proved by executing the following **IS1** and **IS2** for each $P_i(1 \leq i \leq n)$ and each action $t = (sf_i, d_i, sf'_i, C_i, R_i) \in T_X$, where $X \in \{A, B\}$.

IS1 Identify every parameter p and global state $gs \in GS(P_i)$ such that $(\langle d, p \rangle, X)$ is executable in gs .

This is examined by solving the simultaneous inequality consisting of C_i and all AF4 type atoms in P_i .

IS2 For every executable parameter n obtained by IS1, show

$$\forall gs \in GS(P_i) \forall gs'(gs \xrightarrow{X^{(d_i, n)}} gs') \exists j(1 \leq j \leq n) \{gs' \in GS(P_j)\}.$$

A verification system based on this method was developed (see [9].)

4. A Method for Verifying Liveness

4.1 Q-liveness

We formulate liveness property as Q -liveness defined as below.

Definition 4 For a protocol Π , let GS_Q be the set of global states which satisfy a property Q . If $\forall gs \in RS_\Pi \exists gs' \in GS_Q \{gs \xrightarrow{*} gs'\}$ holds, the protocol Π is said to be Q -live, where RS_Π denotes the reachable set of Π . $\square$

Table 4: A logical formula E_{EX} for the protocol Π_{EX} in 3.3

$P_1 = \langle \text{STA713}_A, \text{STA713}_B \rangle$ $\wedge \langle \varepsilon, \varepsilon \rangle$ $\wedge VM_A = VM_B$ $\wedge VA_A = VA_B$ $\wedge VM_A = VA_A$ $\wedge VM_B = VA_B;$	$P_2 = \langle \text{STA713}_A, \text{STA713}_B \rangle$ $\wedge \langle \varepsilon, \varepsilon \rangle$ $\wedge VM_A = VM_B$ $\wedge VA_A = VA_B$ $\wedge VM_A > VA_A$ $\wedge VM_B > VA_B;$
$P_3 = \langle \text{STA713}_A, \text{STA713}_B \rangle$ $\wedge \langle \varepsilon, \text{MIP} \cdot \text{MIP}^* \rangle$ $\wedge \text{step1}(u_{AB})$ $\wedge VM_A = \text{lastp}(u_{AB}) + 1$ $\wedge VM_B = \text{firstp}(u_{AB})$ $\wedge VA_A = VA_B$ $\wedge VM_A > VA_A$ $\wedge VM_B = VA_B;$	$P_4 = \langle \text{STA713}_A, \text{STA713}_B \rangle$ $\wedge \langle \varepsilon, \text{MIP} \cdot \text{MIP}^* \rangle$ $\wedge \text{step1}(u_{AB})$ $\wedge VM_A = \text{lastp}(u_{AB}) + 1$ $\wedge VM_B = \text{firstp}(u_{AB})$ $\wedge VA_A = VA_B$ $\wedge VM_A > VA_A$ $\wedge VM_B > VA_B;$
$P_5 = \langle \text{STA713}_A, \text{STA713}_B \rangle$ $\wedge \langle \text{MIA} \cdot \text{MIA}^*, \varepsilon \rangle$ $\wedge \text{st-inc}(u_{BA})$ $\wedge VM_A = VM_B$ $\wedge VA_B = \text{lastp}(u_{BA}) + 1$ $\wedge VA_A \leq \text{firstp}(u_{BA})$ $\wedge VM_A > VA_A$ $\wedge VM_B = VA_B;$	$P_6 = \langle \text{STA713}_A, \text{STA713}_B \rangle$ $\wedge \langle \text{MIA} \cdot \text{MIA}^*, \varepsilon \rangle$ $\wedge \text{st-inc}(u_{BA})$ $\wedge VM_A = VM_B$ $\wedge VA_B = \text{lastp}(u_{BA}) + 1$ $\wedge VA_A \leq \text{firstp}(u_{BA})$ $\wedge VM_A > VA_A$ $\wedge VM_B > VA_B;$
$P_7 = \langle \text{STA713}_A, \text{STA713}_B \rangle$ $\wedge \langle \text{MIA} \cdot \text{MIA}^*, \text{MIP} \cdot \text{MIP}^* \rangle$ $\wedge \text{step1}(u_{AB})$ $\wedge \text{st-inc}(u_{BA})$ $\wedge VM_A = \text{lastp}(u_{AB}) + 1$ $\wedge VM_B = \text{firstp}(u_{AB})$ $\wedge VA_B = \text{lastp}(u_{BA}) + 1$ $\wedge VA_A \leq \text{firstp}(u_{BA})$ $\wedge VM_A > VA_A$ $\wedge VM_B = VA_B;$	$P_8 = \langle \text{STA713}_A, \text{STA713}_B \rangle$ $\wedge \langle \text{MIA} \cdot \text{MIA}^*, \text{MIP} \cdot \text{MIP}^* \rangle$ $\wedge \text{step1}(u_{AB})$ $\wedge \text{st-inc}(u_{BA})$ $\wedge VM_A = \text{lastp}(u_{AB}) + 1$ $\wedge VM_B = \text{firstp}(u_{AB})$ $\wedge VA_B = \text{lastp}(u_{BA}) + 1$ $\wedge VA_A \leq \text{firstp}(u_{BA})$ $\wedge VM_A > VA_A$ $\wedge VM_B > VA_B;$

For example, if Q states that some interested action t is executable, then Q -liveness means that the action t is eventually executed from any reachable state. If Q characterizes normal completion, then Q -liveness means that the protocol can always normally complete.

For a protocol Π , a directed graph $RG_\Pi = (RS_\Pi, \{(gs, gs') \mid gs \rightarrow gs'\})$ is called the *reachability graph* of Π , where RS_Π denotes the reachability set of Π . If the reachability set RS_Π is finite, RG_Π is also finite. In such case, Q -liveness can be verified by exploring RG_Π for a property Q described as temporal logic formula. Such a verification method is

called a model checking algorithm^[4]. However, if the reachability set RS_Π is infinite, it is impossible to apply the model checking algorithm for verifying Q -liveness of the protocol Π .

In the rest of this paper, we consider a method for verifying Q -liveness of 2-ECFSMs in which RS_Π is infinite. Let the reachability set RS_Π be divided into finite numbers of distinguished subsets $GS_1, GS_2, \dots, GS_n$. Let $V_F = \{v_1, v_2, \dots, v_n\}$, where n is the number of distinguished subsets introduced above. Let $G = (V_F, E)$ be a directed graph satisfying the following, where “ $\Rightarrow$ ” denotes an implication:

$$(v_i, v_j) \in E \Rightarrow \forall gs \in GS_i \exists gs' \in GS_j \{gs \rightarrow gs'\}. \quad (*2)$$

Let $GS_Q = GS_i$ for some i ($1 \leq i \leq n$), it is considerable to verify Q -liveness of Π by exploring G . However, since the combination of “ $\forall gs \in GS_i$ ” and direct reachability in (*2) is too strict, it seems that most executable transitions of given protocol may not be represented in such graph. Therefore, it is often the case that such a graph does not preserve the characteristics of given protocol.

Definition 5 For a protocol Π and an invariant formula $F = P_1 \vee P_2 \vee \dots \vee P_n$, a directed graph $G = (V_F = \{v_i \mid 1 \leq i \leq n\}, E)$ which satisfies the following is called a *degenerated reachability graph* of Π :

$$\{\text{there exists a path from } v_i \text{ to } v_j \text{ in } G\} \Rightarrow \forall gs \in GS_i \exists gs' \in GS_j \{gs \xrightarrow{*} gs'\}. \quad \square$$

4.2 A Degenerated Reachability Graph and Q -liveness

Let $F = P_1 \vee P_2 \vee \dots \vee P_n$ be an invariant of a protocol Π and assume that $GS(F)$ contains no unsafe state. We will consider the verification of Q -liveness, where Q is specified as $GS_Q = \bigcup_{i \in I_Q} GS(P_i)$, where I_Q is a subset of the indexes $\{1 \leq i \leq n\}$ of the disjuncts P_i of the invariant F .

Since $GS_Q = \bigcup_{i \in I_Q} GS(P_i)$, if the following formula holds, the Q -liveness of a protocol Π holds:

$$\forall i(1 \leq i \leq n) \forall gs \in GS(P_i) \exists j \in I_Q \exists gs' \in GS(P_j) \{gs \xrightarrow{*} gs'\}. \quad (*3)$$

In the following, if gs' is directly reachable from gs by a sending event of PM_A (or PM_B), we denote $gs \xrightarrow{A \cdot s} gs'$ (or $gs \xrightarrow{B \cdot s} gs'$). If gs' is directly reachable from gs by a receiving event of PM_A (or PM_B), we denote $gs \xrightarrow{A \cdot r} gs'$ (or $gs \xrightarrow{B \cdot r} gs'$). The transitive closure of the relations “ $\xrightarrow{A \cdot r}$ ” and “ $\xrightarrow{B \cdot r}$ ” are denoted by “ $\xrightarrow{A \cdot r}^+$ ” and “ $\xrightarrow{B \cdot r}^+$ ” respectively.

As a degenerated reachability graph of Π , we will consider the following directed graph $DRG = (V_F, E_S \cup E_R^+)$:

$$\begin{aligned} V_F &= \{v_i \mid 1 \leq i \leq n\}; \\ E_S &= \{(v_i, v_j) \mid \forall gs \in GS(P_i) \exists gs' \in GS(P_j) \{gs \xrightarrow{A \cdot s} gs'\}\} \\ &\quad \cup \{(v_i, v_j) \mid \forall gs \in GS(P_i) \exists gs' \in GS(P_j) \{gs \xrightarrow{B \cdot s} gs'\}\}; \\ E_R^+ &= \{(v_i, v_j) \mid \forall gs \in GS(P_i) \exists gs' \in GS(P_j) \{gs \xrightarrow{A \cdot r}^+ gs'\}\} \\ &\quad \cup \{(v_i, v_j) \mid \forall gs \in GS(P_i) \exists gs' \in GS(P_j) \{gs \xrightarrow{B \cdot r}^+ gs'\}\}. \end{aligned}$$

If there exists a path from v_i to v_j in DRG , the following formula holds:

$$\forall gs \in GS(P_i) \exists gs' \in GS(P_j) \{gs \xrightarrow{*} gs'\}.$$

Since F is an invariant in Π , if the following formula holds, then (*3) holds:

$$\forall i(1 \leq i \leq n) \exists j \in I_Q \{ \text{there exists a path from } v_i \text{ to } v_j \text{ in } DRG \}. \quad (*4)$$

4.3 Constructing DRG

In this subsection, we will describe a method for constructing a degenerated reachability graph $DRG = (V_F, E_S \cup E_R^+)$.

4.3.1 Constructing E_S

For P_i, P_j ($1 \leq i, j \leq n$) and a sending action $t = (sf, d, sf', C, R)$ of PM_X ($X \in \{A, B\}$), if the following conditions (a) and (b) hold, then $(v_i, v_j) \in E_S$:

- (a) For every $gs = (s_A, s_B, ch_{BA}, ch_{AB}) \in GS(P_i)$, there exists a parameter value p such that $\langle d, p \rangle$ is executable in s_X . This condition is written in the following (ES1)–(ES3), where C_{P_i} is a simultaneous inequality consisting of all AF4 type atoms in P_i . The set of inequalities in C is divided into C_p and $C_{\bar{p}}$, where C_p consists of inequalities that contain the parameter variable p $C_{\bar{p}}$ consists of inequalities that does not contain p :

(ES1) $sf_X(P_i) = sf$;

(ES2) if the register values of PM_X $\langle p_1, p_2, \dots, p_r \rangle$ satisfy C_{P_i} then $\langle p_1, p_2, \dots, p_r \rangle$ also satisfies $C_{\bar{p}}$;

(ES3) if the register values of PM_X $\langle p_1, p_2, \dots, p_r \rangle$ satisfy C_{P_i} , then there exists a parameter value p such that $\langle p_1, p_2, \dots, p_r, p \rangle$ satisfy C_p and C_{P_i} .

(b)

$$\forall gs \in GS(P_i), p \forall gs' (gs \xrightarrow{X}^{(d, p)} gs') \{ gs' \in GS(P_j) \}. \quad (*5)$$

(*5) can be verified in a similar way to a part of the procedure described in [9]. (ES1) can be immediately verified from the values of t and P_i . The procedure for verifying (ES2) and (ES3) will be described in 4.4. In the above way, several edges can be verified to be in E_S . Let $E'_S \subseteq E_S$ be the set of such edges.

4.3.2 Constructing E_R^+

In the following discussion, for the AF2 type atom $\langle R_{BA}, R_{AB} \rangle$ in every P_i ($1 \leq i \leq n$), we assume that R_{BA} (or R_{AB}) specifies either a set consisting of only the empty sequence, that is $R_{BA} = \varepsilon$ (or $R_{AB} = \varepsilon$), or a set which does not contain the empty sequence.

We will consider the following graphs $G_{A,R} = (V_F, E_{A,R})$ and $G_{B,R} = (V_F, E_{B,R})$ whose edges express possibilities of transitions between two sets of global states by a reception by a specified protocol machine:

$$E_{A,R} = \{ (v_i, v_j) \mid \exists gs \in GS(P_i) \exists gs' \in GS(P_j) \{ gs \xrightarrow{A, \tau} gs' \} \};$$

$$E_{B,R} = \{ (v_i, v_j) \mid \exists gs \in GS(P_i) \exists gs' \in GS(P_j) \{ gs \xrightarrow{B, \tau} gs' \} \}.$$

For every P_i, P_j ($1 \leq i, j \leq n$), $(v_i, v_j) \in E_{A,R}$ (or $(v_i, v_j) \in E_{B,R}$) iff there exists a receiving action $t = (sf, d, sf', C, R)$ of PM_A (or PM_B) satisfying the following (a) and (b):

- (a) There exists a global state $gs = (s_A, s_B, u_{BA}, u_{AB}) \in GS(P_i)$ and a parameter value p such that $\langle d, p \rangle$ is executable, i.e. the following (ER1)–(ER3) holds, where C_{P_i} is the simultaneous inequality consisting of all AF4 type atoms in P_i and $\langle R_{BA,i}, R_{AB,i} \rangle$ is the AF2 type atom in P_i :

(ER1) $sf_A(P_i) = sf$ (or $sf_B(P_i) = sf$);

(ER2) $\exists u \in \Sigma_{BA}^* \{d \cdot u \in L(R_{BA,i})\}$ (or $\exists u \in \Sigma_{AB}^* \{d \cdot u \in L(R_{AB,i})\}$);

(ER3) there exists register values $p_1, p_2, \dots, p_r$ of PM_A (or PM_B) and a parameter value p of the receiving message $\langle d, p \rangle$ satisfying the simultaneous inequality consisting of C_{P_i} and C .

(b)

$$\exists gs \in GS(P_i), p, gs' \in GS(P_j) \{gs \xrightarrow{A(\text{or } B)}^{(d,p)} gs'\}. \quad (*6)$$

(ER1) can be immediately determined from the values of t and P_i . (ER2) and a necessary condition $(*6)'$ of $(*6)$ also can be determined in a similar way to a part of the procedure described in [9]. The procedure for verifying (ER3) will be described in 4.4. We shall let $E'_{A,R}$ (or $E'_{B,R}$) be a set of edges (v_i, v_j) such that (ER1), (ER2), (ER3), and $(*6)'$ holds for some action t .

Let $\langle R_{BA,i}, R_{AB,i} \rangle$ be an AF2 type atom in P_i , consider the following the subsets of the node set V_F of DRG :

$$V_{BA=\varepsilon} = \{v_i \mid R_{BA,i} = \varepsilon\};$$

$$V_{AB=\varepsilon} = \{v_i \mid R_{AB,i} = \varepsilon\}.$$

Since we assumed that $GS(F)$ does not contain any unspecified reception state, for every global state gs , there exists a global state gs' reachable from gs such that the communication channel from PM_B to PM_A is empty in gs' , i.e. the following condition holds:

$$\forall gs \in GS(P_i) \exists gs' \in \bigcup_{v_j \in V_{BA=\varepsilon}} GS(P_j) \{gs \xrightarrow{A \cdot T}^+ gs'\}. \quad (*7)$$

Lemma 1 An edge (v_i, v_k) is in E_R^+ , if a node v_k occurs on every path from v_i to every $v_j \in V_{BA=\varepsilon}$ in $G'_{A,R} = (V_F, E_{A,R})$.

[Proof] Assume that a node v_k occurs on every path from v_i to every $v_j \in V_{BA=\varepsilon}$ in $G'_{A,R}$. Then, the following formula holds:

$$\forall gs \in GS(P_i) \{ \exists gs' \in \bigcup_{v_j \in V_{BA=\varepsilon}} GS(P_j) \{gs \xrightarrow{A \cdot T}^+ gs'\} \Rightarrow \exists gs' \in GS(P_k) \{gs \xrightarrow{A \cdot T}^+ gs'\} \}. \quad (*8)$$

By $(*7)$ and $(*8)$, $\forall gs \in GS(P_i) \exists gs' \in GS(P_k) \{gs \xrightarrow{A \cdot T}^+ gs'\}$ holds. Therefore, $(v_i, v_k) \in E_R^+$. $\square$

We can obtain the same lemma for $E'_{B,R}$.

By checking the sufficient condition in Lemma 1 for every pair of v_i and v_j , several edges can be verified to be in E_R^+ . Let $E'_R \subseteq E_R^+$ be a set of such edges.

Figure 1: DRG_{Π} for Π_{EX} and F_{EX} in 3.3

4.3.3 Exploring DRG

Since $E'_S \subseteq E_S$, $E'_R \subseteq E_R^+$, if there exists a path from v_i to v_j in $(V_F, E'_S \cup E'_R)$, there also exists a path from v_i to v_j in $DRG = (V_F, E_S \cup E_R^+)$. Therefore, if the following holds, (*4) and Q -liveness holds:

$$\forall i(1 \leq i \leq n) \exists j \in I_Q \{ \text{there exists a path from } v_i \text{ to } v_j \}.$$

Figure 1 shows $(V_F, E'_S \cup E'_R)$ obtained by the above procedure on the protocol Π_{EX} and invariant F_{EX} in 3.3. Since there exists a path from every node to v_i ($1 \leq i \leq n$) in $(V_F, E'_S \cup E'_R)$, Π_{EX} is P_i -live for every i ($1 \leq i \leq n$).

4.4 Evaluating Presburger Formula

Let C be a simultaneous linear inequality with m variables $x_1, x_2, \dots, x_m$, c be a linear inequality whose variables are common with those of C , and C_y be a simultaneous linear inequality whose variables are common with those of C except only one variable y . The decision problems (ER3), (ES2) and (ES3) in 4.3 can be written in the following (S1), (S2), (S3) respectively, where X denotes $\langle x_1, x_2, \dots, x_m \rangle$:

- (S1) There exists an $S \in \mathcal{N}^m$ such that $X = S$ is a solution of C ;
- (S2) If $X = S(\in \mathcal{N}^m)$ is a solution of C , then $X = S$ satisfies c ;
- (S3) If $X = S(\in \mathcal{N}^m)$ is a solution of C , then there exists an $s \in \mathcal{N}$ such that $(X, y) = (S, s)$ is a solution of a simultaneous inequality consisting of C and C_y .

These problems can be expressed as formulae of Presburger arithmetic, which is known to be decidable^[5]. However, (S1) is equivalent to the general form of integer linear programming, which is known to be $\mathcal{NP}$ -complete^[5].

In the following, we consider a special case such that (S1), (S2) and (S3) can be efficiently solved.

Definition 6 A linear inequality of the following form is called a *difference constraint*, where x and y are variables, and q is a nonnegative integer:

$$x - y \leq q.$$

□

If every AF4 type atom in F and every inequality in the action condition C of every action of PM_A or PM_B are difference constraints, then every inequality in (EA3), (ES2), (ES3) is also a difference constraint. For the satisfiability problem of a simultaneous inequality consisting of difference constraints, which is called a system of difference constraints, the following method is known^[11].

4.4.1 Solving a System of Difference Constraints

For a system of difference constraints C with m variables, construct the weighted digraph $G_C = (V_C, E_C)$ as follows:

- $V_C = \{v_1, v_2, \dots, v_m\}$;
- $E_C = \{(v_i, v_j) \mid x_j - x_i \leq q \text{ is a difference constraint in } C\}$;
- For a difference constraint $x_j - x_i \leq q$ in C , the weight $w(v_i, v_j)$ of the edges (v_i, v_j) is q

Then the following lemmas hold^[11].

Lemma 2 A system of difference constraints has a solution $S \in \mathcal{N}^m$ iff G_C contains no negative-weight cycles. □

Lemma 3 $x_j - x_i$ is bounded under the system of difference constraints C iff G_C contains a path from v_i to v_j . If G_C contains a path from v_i to v_j , let $d(v_i, v_j)$ be the weight of the shortest path from v_i to v_j , then the following are equivalent:

- There exists a solution of C such that $x_j = s_j, x_i = s_i$;
- $s_j - s_i \leq d(v_i, v_j)$.

□

Bellman-Ford algorithm determines in $O(|V_C| \cdot |E_C|)$ whether a given G_C contains no negative-weight cycles^[11]. It also finds a shortest path and its weight of a given G_C for any pair of source node and destination node.

4.4.2 Evaluating a Class of Presburger Formula

In the following of this section, we suppose that every inequality in (S1)–(S3) is a difference constraint and let l and l_y be the numbers of inequalities in C and C_y respectively.

Obviously, (S1) can be solved in $O(lm)$. Let the constraint c in (S2) be “ $x_j - x_i \leq q$ ”. By determining whether there exists no solution for the system of difference constraints consisting of C and the constraint “ $x_i - x_j \leq -q - 1$ ”, which is the negation of c , (S2) can be solved in $O(lm)$.

Let C_{-y} and C_{+y} be the set of difference constraints in C_y such that the coefficient of y are “ -1 ” and “ $+1$ ” respectively. For $c' \in C_{-y}$ and $c'' \in C_{+y}$, let $E(C, c', c'')$ be the predicate which states the following:

$$\forall S(\text{solution of } C) \exists s_y \{(S, s_y) \text{ satisfies } c' \text{ and } c''\}.$$

Lemma 4 (S3) holds iff $E(C, c', c'')$ holds for every pair of $c' \in C_{-y}$, and $c'' \in C_{+y}$.

[**Proof**] The necessity is trivial. To show the sufficiency, suppose that (S3) does not hold, then there exists a solution S of C such that:

$$\neg \exists s_y \{ (S, s_y) \text{ satisfies } C_y \}. \quad (*9)$$

Let $S = (s_1, s_2, \dots, s_m)$ be such a solution. By assigning S to X , every constraint $s_j - y \leq q$ in C_{-y} gives a lower bound $s_j - q$ of y . Let c' be the strongest constraint among them. Similarly, by assigning S to X , every constraint $y - s_i \leq q$ in C_{+y} gives an upper bound $q + s_i$ of y . Let c'' be the strongest constraint among them. Then, (*9) is equivalent to the followings:

$$\neg \exists s_y \{ (S, s_y) \text{ satisfies } c' \text{ and } c'' \}. \quad (*10)$$

Therefore, $E(C, c', c'')$ does not hold. $\square$

The above lemma states that (a) if there exists a solution of C , (S3) can be determined by determining $E(C, c', c'')$ for every pair of $c' \in C_{-y}$ and $c'' \in C_{+y}$, (b) otherwise (S3) vacuously holds.

Let c' be $x_j - y \leq q'$ and c'' be $y - x_i \leq q''$.

Lemma 5 $E(C, c', c'')$ holds iff there exists a path from v_i to v_j in G_C and $d(v_i, v_j) \leq q' + q''$, where $d(v_i, v_j)$ is the weight of a shortest path from v_i to v_j .

[**Proof**] Clearly, the following relation holds.

$$x_j - x_i \leq q' + q'' \Leftrightarrow \exists y \{ x_j - y \leq q' \wedge y - x_i \leq q'' \}. \quad (*11)$$

Suppose that G_C does not contain a path from v_i to v_j . It follows from Lemma 3 that there exists a solution S of C such that $s_j - s_i > q' + q''$. By (*11) we can see that $E(C, c', c'')$ is false.

Suppose that $d(v_i, v_j) > q' + q''$. Consider s_j and s_i such that $d(v_i, v_j) \geq s_j - s_i > q' + q''$. By (*11), there exists no s_y such that $s_j - s_y \leq q'$ and $s_y - s_i \leq q''$. On the other hand, it follows from Lemma 3 that there exists a solution of C such that $x_j = s_j$ and $x_i = s_i$. Therefore, $E(C, c', c'')$ does not hold.

Suppose that $d(v_i, v_j) \leq q' + q''$. It follows from Lemma 3 that $s_j - s_i \leq q' + q''$ holds for each solution $\langle s_1, s_2, \dots, s_m \rangle$ of C . By (*11), $E(C, c', c'')$ holds. $\square$

Since, the Bellman-Ford algorithm can determine the existence of a path between two nodes and calculate the weight of shortest paths between two nodes, $E(C, c', c'')$ can be determined in $O(lm)$. Consequently, the following theorem holds.

Theorem 1 (S3) can be determined in $O(lm(l_y)^2)$. $\square$

5. An Experimental Result

To show the usefulness of the proposed method, we performed an experiment to verify liveness of a sample protocol using a verification system based on the above method.

5.1 A Verification System

We developed a verification system by incorporating the procedures described in section 4 to the verification system described in [9].

An input to the verification system consists of the definition of protocol machines, properties of predicates on sequence of integers such as *step1* explained in Section 3.2 and Table 3, a logical formula $F = P_1 \vee P_2 \vee \dots \vee P_n$, and I_Q which defines the property Q . A human verifier can use macro notations to describe logical formulae.

The verification system executes the procedures described in [9] and Section 4 to verify safety and liveness of a given protocol. The system also includes the efficient procedure described in 4.4 to determine a restricted form of Presburger formula.

The system is described in C, lex and yacc, and is executed on UNIX environment. The size of the source code of this system is about 16,300 lines.

5.2 Extracting a Sample Protocol

As a sample protocol of non-channel-bounded 2-ECFSMs, $\Pi_{\text{ses}} = (PM_{\text{ses}_A}, PM_{\text{ses}_B})$ was extracted from the *data transfer phase* of *kernel*, *duplex*, *minor synchronize* and *major synchronize* functional units from the OSI session protocol. For simplification, two tokens *ma* and *mi* are integrated into a token *ma-mi*.

For the protocol Π_{ses} , two protocol machines PM_{ses_A} and PM_{ses_B} are isomorphic concerning to S , Σ , and T but differ in those initial states, i.e. PM_{ses_A} owns *ma-mi* token while PM_{ses_B} does not, at those initial states. The size of the states of finite controls and the numbers of registers of the protocol machines are 10 and 2 respectively. The number of message types used in the protocol is 12. The number of actions of each protocol machine is 22. In the definition of the actions of the machines, the action condition C of every action can be given in the form of difference constraints.

5.3 An Experimental Result

At first, a human verifier divided a set of global states expected to be reachable into 144 subsets and described a logical formula F_{SAFE} ¹. The verification system showed that F_{SAFE} is invariant of Π_{ses} and $GS(F_{\text{SAFE}})$ does not contain any unsafe state in Π_{ses} . Hence, it can be concluded that Π_{ses} is safe.

For verifying liveness, a human verifier divided $GS(P_i)$ for each disjunct P_i of F_{SAFE} into several distinguished subsets by the following criteria:

- Is communication channel from PM_{ses_B} to PM_{ses_A} empty?;
- Is communication channel from PM_{ses_A} to PM_{ses_B} empty?;
- Does a condition for executing of each action hold?

Finally, the human verifier divided the reachability set of Π into 480 subsets, and described a formula F_{LIVE} based on those subsets. The numbers of atomic formulae of each types AF1–AF4 in F_{LIVE} are 30, 30, 46, 471 respectively. The macro processor unfolds F_{LIVE} and obtains a disjunctive normal form F'_{LIVE} (equivalent to F_{LIVE}) consisting of 480 disjuncts. Both in F_{SAFE} and F_{LIVE} , every AF4 type atom specified as difference

¹For describing F_{SAFE} or F_{LIVE} , a human verifier used 2 predicates and 2 defined functions, which are similar to those of [9].

constraint. Since each action condition is also specified as difference constraints, every (ER3), (ES2) and (ES3) are determined by the procedure discussed in 4.4

At first, the verification system showed that F'_{LIVE} is an invariant of Π_{ses} and $GS(F'_{LIVE})$ does not contain any unsafe state in Π_{ses} . The verification system also showed P_i -liveness for every disjunct P_i of F'_{LIVE} . During the procedure of verifying liveness, the Bellman-Ford procedure were called 27,628 times. For the degenerated reachability graph $DRG = (V'_{LIVE}, E_S \cup E_R^+)$ constructed in the procedure of verifying liveness, $|V'_{LIVE}| = 480$, $|E_S| = 1,566$, and $|E_R^+| = 1,300$.

The above experiments were executed on a UNIX workstation(NWS-5000, 64MB). The CPU time in the executions of verifying safety and liveness were 40 seconds and 60 seconds respectively.

6. Conclusion

In this paper, we formulated liveness property as Q -liveness and propose a method for verifying Q -liveness of communication protocol modeled as 2-ECFSMs. We also described a verification system based on the proposed method and a verification example extracted from OSI session protocol.

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