

AD-A171 536

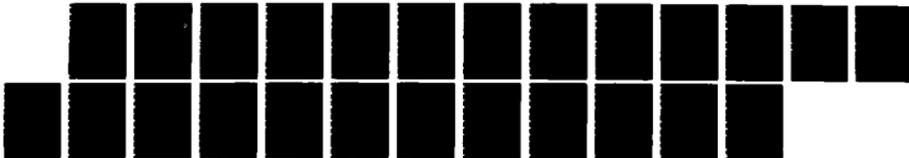
THE POWER OF THE QUEUE(U) MASSACHUSETTS INST OF TECH  
CAMBRIDGE LAB FOR COMPUTER SCIENCE M LI ET AL. APR 86  
MIT/LCS/TM-303 N00014-83-K-0125

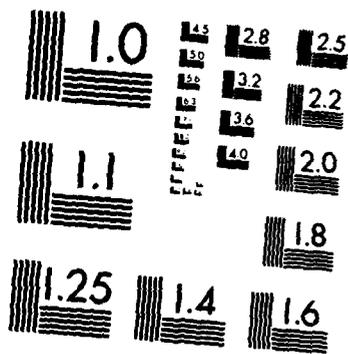
1/1

UNCLASSIFIED

F/G 9/2

NL

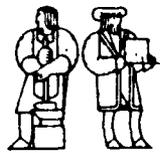




MICROCOPY RESOLUTION TEST CHART  
NATIONAL BUREAU OF STANDARDS-1963-A

12

LABORATORY FOR  
COMPUTER SCIENCE



MASSACHUSETTS  
INSTITUTE OF  
TECHNOLOGY

AD-A171 536

MIT/LCS/TM-303

DTIC  
COLLECTED  
SEP 03 1986  
S D

THE POWER OF THE QUEUE

MING LI  
LUC LONGPRE  
PAUL M. B. VITANYI

THIS FILE COPY

APRIL 1986

86 9 02 129

57 TECHNOLOGY SQUARE, CAMBRIDGE, MASSACHUSETTS 02139

**DISTRIBUTION STATEMENT A**  
Approved for public release  
Distribution Unlimited

REPORT DOCUMENTATION PAGE		READ INSTRUCTIONS BEFORE COMPLETING FORM
1. REPORT NUMBER MIT/LCS/TM-303	2. GOVT ACCESSION NO. <b>ADA171536</b>	3. RECIPIENT'S CATALOG NUMBER
4. TITLE (and Subtitle) The power of the queue.		5. TYPE OF REPORT & PERIOD COVERED Interim research April 1986
		6. PERFORMING ORG. REPORT NUMBER MIT/LCS/TM-303
7. AUTHOR(s) Ming Li, Luc Longpre, and Paul M.B. Vitanyi		8. CONTRACT OR GRANT NUMBER(s) DARPA/DOD N00014-83-K-0125
9. PERFORMING ORGANIZATION NAME AND ADDRESS MIT Laboratory for Computer Science 545 Technology Square Cambridge, MA 02139		10. PROGRAM ELEMENT, PROJECT, TASK AREA & WORK UNIT NUMBERS
11. CONTROLLING OFFICE NAME AND ADDRESS DARPA/DOD 1400 Wilson Boulevard Arlington, VA 22209		12. REPORT DATE April 1986
		13. NUMBER OF PAGES 22
14. MONITORING AGENCY NAME & ADDRESS (if different from Controlling Office) ONR/Department of the Navy Information Systems Program Arlington, VA 22217		15. SECURITY CLASS. (of this report) Unclassified
		15a. DECLASSIFICATION/DOWNGRADING SCHEDULE
16. DISTRIBUTION STATEMENT (of this Report) Approved for public release, distribution is unlimited		
17. DISTRIBUTION STATEMENT (of the abstract entered in Block 20, if different from Report) Unlimited		
18. SUPPLEMENTARY NOTES		
19. KEY WORDS (Continue on reverse side if necessary and identify by block number) tape, stack, queue, pushdown stores, determinism, nondeterminism, off-line one-way input, time complexity, lower bound, upper bound, simulation, algorithmic information theory, Kolmogorov complexity		
20. ABSTRACT (Continue on reverse side if necessary and identify by block number) Queues, stacks (pushdown stores), and tapes are storage models which have direct applications in compiler design and the general design of algorithms. Whereas stacks (pushdown store or last-in-first-out storage) have been thoroughly investigated and are well understood, this is much less the case for queues (first-in-first-out storage). This paper contains a comprehensive study comparing queues to stacks and tapes. We address off-line		

machines with a one-way input, both deterministic and nondeterministic. The techniques rely on algorithmic information theory (Kolmogorov Complexity).

- P. -

## The Power of the Queue

*Ming Li*

Department of Computer and Information Science  
The Ohio State University, Columbus, OH 43210

*Luc Longpré*

Department of Computer Science  
University of Washington, Seattle, WA 98195

*Paul M.B. Vitányi*

M.I.T. Laboratory for Computer Science  
Cambridge, MA 02139  
and  
Centrum voor Wiskunde en Informatica  
Amsterdam, The Netherlands

*April, 1986*

### ABSTRACT

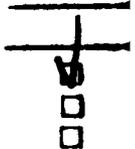
Queues, stacks (pushdown stores), and tapes are storage models which have direct applications in compiler design and the general design of algorithms. Whereas stacks (pushdown store or last-in-first-out storage) have been thoroughly investigated and are well understood, this is much less the case for queues (first-in-first-out storage). This paper contains a comprehensive study comparing queues to stacks and tapes. We address off-line machines with a one-way input, both deterministic and nondeterministic. The techniques rely on algorithmic information theory (Kolmogorov Complexity).

*Keywords & Phrases* : tape, stack, queue, pushdown stores, determinism, nondeterminism, off-line one-way input, time complexity, lower bound, upper bound, simulation, algorithmic information theory, Kolmogorov complexity

*Note* : This paper will be presented at the *Structure in Complexity Theory Conference*, to be held at the University of California, Berkeley, June 2-5, 1986.

© 1986 Massachusetts Institute of Technology, Cambridge, MA 02139

The work of the third author was supported in part by the Office of Naval Research under Contract N00014-85-K-0168, by the Office of Army Research under Contract DAAG29-84-K-0058, by the National Science Foundation under Grant DCR-83-02391, and by the Defense Advanced Research Projects Agency (DARPA) under Grant N00014-83-K-0125



Codes

Dist	Special
A-1	

## 1. Introduction

Queues, stacks (pushdown stores), and tapes are storage models which have direct applications in compiler design and the general design of algorithms. Whereas stacks (pushdown store or last-in-first-out storage) have been thoroughly investigated and are well understood, this is much less the case for queues (first-in-first-out storage). In this paper we present a comprehensive study comparing queues to stacks and tapes. We address off-line machines with a one-way input. In particular, 1 queue and 1 tape (or stack) are not comparable:

- (1) Simulating 1 stack (and hence 1 tape) by 1 queue requires  $\Omega(n^{4/3}/\log n)$  time in both the deterministic and the nondeterministic cases.
- (2) Simulating 1 queue by 1 tape requires  $\Omega(n^2)$  time in the deterministic case, and  $\Omega(n^{4/3}/\log n)$  in the nondeterministic case;

We further compare the relative power between different numbers of queues:

- (3) Nondeterministically simulating 2 queues (or 2 tapes) by 1 queue requires  $\Omega(n^2/(\log^2 n \log \log n))$  time and deterministically simulating 2 queues (or 2 tapes) by 1 queue requires  $\Omega(n^2)$  time. The second bound is tight. The first is almost tight.
- (4) We also obtain the simulation results for queues: 2 nondeterministic queues (or 3 pushdown stores) can simulate  $k$  queues in linear time. One queue can simulate  $k$  queues in quadratic time.

It has been known for over twenty years that all multitape Turing machines can be simulated on-line by 2-tape Turing machines in time  $O(n \log n)$  [HS2], and by 1-tape Turing machines in time  $O(n^2)$  [HU]. Since then, many other models of computation have been introduced and compared. (See [Aa, DGPR, HS1, HS2, HU, LS, PSS, Pa, Vi2].) In addition to different storage mechanisms, real-time, on-line and off-line machines have been studied. An on-line machine is expected to give an answer after reading each prefix of the input. In this paper, we consider the off-line machines, where an answer is given only once the whole input has been read. We also use the one-way input convention, where the machine has a one-way input, a finite control and access to some storage.

The relative power of stacks and tapes is more or less well known. For example, for the nondeterministic case, we know that 1 stack  $<$  1 tape  $<$  2 stacks  $<$  3 stacks  $=$   $k$  stacks  $=$   $k$  tapes, where  $A < B$  means that  $B$  can simulate  $A$  in linear time, while  $A$  cannot simulate  $B$  in linear time. In most of the cases, close lower and upper bounds are known for the simulation [Ma, Li1, Vi1, LV, Li2].

In this paper, we give a complete characterization of (off-line) queue machines. The main theorems show that one queue machines are not comparable to one stack or one tape machines, both deterministically and nondeterministically. We also compare the relative power of machines having different number of queues. We use Kolmogorov complexity techniques [Ko, Ch. So] to prove the theorems, together with some new techniques to enable us to deal with queues. The Kolmogorov complexity of a string  $x$ ,  $K(x)$ , is the length of the shortest program printing the string  $x$ . By a simple counting

argument, we know that there are strings  $x$  of each length such that  $K(x) \geq |x|$ . These strings are called *incompressible* or *K-random*.

In section 2, we introduce the jamming lemma which is used in further proofs. In section 3, we show that deterministically simulating a queue by a tape takes quadratic time (infinitely often). (For the lower bound on the simulation time of 1 queue by 1 tape in the nondeterministic case, see [Li3, LV].) In section 4, we have a lower bound for nondeterministically simulating a stack by a queue. In section 5, we present lower and upper bounds for simulating  $k$  queues by 2 queues or 2 queues by one queue.

## 2. The Jamming Lemma

In this section, we are concerned with one-tape and one-queue off-line TM's where the Turing machine has one 1-way input tape in addition to one work tape or one queue, each with one head. We will call the input tape head  $h_1$  and work tape head or queue head  $h_2$ . We say that a poll occurs when the head  $h_1$  moves one cell. At any time  $t$ ,  $h_i(t)$  denotes the position of head  $h_i$  on its tape.

In the following lower bound proofs, the input will be separated into blocks. We will observe the behavior of the machine as the head polls the successive cells in a block. Although the definitions and the Jamming Lemma are stated with respect to one-tape TM's for simplicity, they also apply to one-queue machines where the work tape is replaced by a queue.

**Definition 2.1:** Let  $x_i$  be a block of input, and  $R$  be a tape segment on the storage tape. We say that  $M$  maps  $x_i$  into  $R$  if  $h_2$  never leaves tape segment  $R$  while  $h_1$  is reading  $x_i$ . We say  $M$  maps  $x_i$  onto  $R$  if  $h_2$  traverses the entire tape segment  $R$  without leaving  $R$  while  $h_1$  reads  $x_i$ .

**Definition 2.2:** A crossing sequence (c.s.) associated with the boundary between two contiguous work tape cells is a sequence of ID's of the form  $(M(t), h_1(t))$ , where  $M(t)$  is the state of the machine at time  $t$ , for each time  $t$  when the machine crosses that boundary.

We prove an intuitively straightforward lemma for one-tape machines with one-way input. The lemma states that  $M$  cannot poll too many input symbols, with  $h_2$  located on a given small tape segment bordered by short c.s.'s, without losing some information. Formally:

**Jamming Lemma.** Let the input string start with  $x\# = x_1x_2 \cdots x_k\#$ , with the  $x_i$ 's blocks of equal length  $C$ . Let  $R$  be a segment of  $M$ 's storage tape and let  $l$  be an integer such that  $M$  maps each block  $x_{j_1}, \dots, x_{j_l}$  (of the  $x_i$ 's) into tape segment  $R$ . The contents of the storage tape of  $M$ , at time  $t_\#$  when  $h_1(t_\#) = |x\#|$  and  $h_1(t_\# - 1) = |x|$ , can be reconstructed by using only the blocks  $x_{j_1} \cdots x_{j_{l-1}}$  which remain from  $x_1 \cdots x_k$  after deleting blocks  $x_{j_l}, \dots, x_{j_1}$ , the final contents of  $R$ , the two final c.s.'s on the left and right boundaries of  $R$ , a description of  $M$  and a description of this discussion.

**Remark 2.3:** If we want to give a description of a sequence of different strings of variable length, we use self-delimiting strings, adding  $O(\log n)$  bits for each string of length  $n$ .

**Remark 2.4:** Roughly speaking, if the number of missing bits  $\sum_{j=1}^l |x_j|$  is greater than the number of added description bits then the Jamming Lemma implies that either  $x = x_1 \cdots x_k$  is not incompressible or some information about  $x$  has been lost.

**Proof of the Jamming Lemma.** Let the two positions at the left boundary and the right boundary of  $R$  be  $l_R$  and  $r_R$ , respectively. We now simulate  $M$ . Put the blocks  $x_j$  of  $x_{j_1} \cdots x_{j_{k-1}}$  in their correct positions on the input tape (as indicated by the  $h_1$  values in the left and right c.s.'s). Run  $M$  with  $h_2$  staying to the left of  $R$ . Whenever  $h_2$  reaches point  $l_R$ , the left boundary of  $R$ , we interrupt  $M$  and check whether the current  $ID$  matches the next  $ID$ , say  $ID_i$ , in the c.s. at  $l_R$ . Subsequently, using  $ID_{i+1}$ , we skip the input up to and including  $h_1(t_{i+1})$ , adjust the state of  $M$  to  $M(t_{i+1})$ , and continue running  $M$ . After we have finished left of  $R$ , we do the same thing right of  $R$ . At the end we have determined the appropriate contents of  $M$ 's tape, apart from the contents of  $R$ , at  $t_{\#}$  (i.e., the time when  $h_1$  reaches  $\#$ ). Inscripting  $R$  with its final contents from the reconstruction description gives us  $M$ 's storage tape contents at time  $t_{\#}$ . Notice that although there are many unknown  $x_i$ 's, they are never polled since  $h_1$  skips over them because  $h_2$  never goes into  $R$ .

**Remark 2.5:** If  $M$  is nondeterministic, then we need to rephrase "contents of storage tape" by "legal contents of storage tape", which simply means that some computation path for the same input would create this storage tape contents.

### 3. Lower bound for simulating one queue by one tape

We present a tight lower time bound for deterministic simulation of one queue by one off-line tape with one-way input. (For a lower bound for the nondeterministic case, see [LV] or [Li3].)

**Remark.** Only in this section 3,  $g(n) \in \Omega(f(n))$  means "there is a positive constant  $\delta$  such that  $g(n) \geq \delta f(n)$  infinitely often". Everywhere else the results hold for the stronger variant of  $\Omega$ : "there exist a positive constant  $\delta$  and a positive integer  $n_0$  such that  $g(n) \geq \delta f(n)$  for all  $n \geq n_0$ ".

Let  $\leq_{\text{prefix}}$  mean 'is a prefix of.' Let  $\Sigma = \{0,1\} \times \{0,1,\bar{0},\bar{1},\epsilon\}$ , where  $\epsilon$  denotes the empty string, and consider the words over  $\Sigma$  of the form

$$(a_1, b_1)(a_2, b_2) \cdots (a_n, b_n) \tag{3.1}$$

such that for all  $i$ ,  $1 \leq i \leq n$ ,

$$a_i \in \{0,1\} \text{ and } b_i \in \{0,1,\bar{0},\bar{1},\epsilon\}$$

$$b_1 b_2 \cdots b_i \leq_{\text{prefix}} \hat{a}_1 b_1 \hat{a}_2 b_2 \cdots \hat{a}_i b_i,$$

where for any pair  $(a, b) \in \Sigma$  we define  $\hat{a}$  by

$$\begin{aligned} \hat{a} &= a \text{ if } b = \epsilon \\ \hat{a} &= \bar{a} \text{ if } b \neq \epsilon . \end{aligned}$$

**Remark.** Words of this form constitute the witness language  $L_q$  below, which is real-time acceptable by a queue but which requires  $\Omega(n^2)$  time for acceptance by a tape. Think of the sequence  $a_1 a_2 \cdots a_n$  as the  $n$ -length sequence of bits to be stored consecutively in the queue, and the sequence  $b_1 b_2 \cdots b_n$ , of length  $m$  ( $m \leq n$ ), as the sequence of bits which are consecutively unstored from the queue. (Note, that while  $a_i \neq \epsilon$  for all  $i$ , it is possible that  $b_i = \epsilon$  for some  $i$  ( $1 \leq i \leq n$ ). That is,  $(a_i, b_i)$  specifies that  $\epsilon$  be unstored.) For technical reasons in the proof below, we have to complicate this scheme. All of the prefix of  $a_1 a_2 \cdots a_n$  which has been stored in the queue previously, needs to remain stored in the queue forever. Nonetheless, to force the queue to operate correctly we need to be able to unstore it. To combine both requirements, each pair  $(a_i, b_i)$  causes the queue not only to store  $a_i$  and to unstore  $b_i$  (possibly  $\epsilon$ ), but also to store  $b_i$  anew. Below we show that the scheme of barred and unbarred  $a_i$ 's, related to whether or not the associated 'unstore'  $b_i$ 's are  $\epsilon$  or not, makes it possible to retrieve the complete sequence of  $a_i$ 's, in the order they have been stored originally, from the queue contents at each instant.

The *witness* language  $L_q$  consists of all words satisfying (3.1). To aid intuition, we can view  $L_q$  as the language accepted by a queue  $Q$  as follows:

- Initially,  $Q$  is empty.
- For all  $i \geq 1$ , input command  $(a_i, b_i)$  to  $Q$  is interpreted by  $Q$  as 'if  $b_i = \epsilon$  then append  $a_i$  to the rear else append  $\bar{a}_i$  to the rear; delete  $b_i$  up front; append  $b_i$  to the rear.' (Here 'action1;action2;action3' denotes the sequential execution of action1, action2 and action3.)
- A word  $(a_1, b_1) \cdots (a_n, b_n)$  is accepted if the sequence of successive front items deleted in the actual computation by  $Q$  on this input is the sequence  $b_1, \dots, b_n$ . All other words are rejected.

The properties of words of form (3.1) we need in the sequel are expressed in the following three lemmas.

**Lemma 3.1:** For a word of the form (3.1),  $|\hat{a}_1 b_1 \hat{a}_2 b_2 \cdots \hat{a}_i b_i| - |b_1 b_2 \cdots b_i| = i$  for all  $i$ ,  $1 \leq i \leq n$ .

**Proof:** Obvious.

**Lemma 3.2:** For a word of the form (3.1) we can reconstruct  $a_1 a_2 \cdots a_n$  from the  $n$ -length suffix of  $\hat{a}_1 b_1 \hat{a}_2 b_2 \cdots \hat{a}_n b_n$ .

**Proof:** Let the  $n$ -length suffix be  $x_1 x_2 \cdots x_n$  with  $x_i \in \{0, 1, \bar{0}, \bar{1}\}$  ( $1 \leq i \leq n$ ). By (3.1) one of the following two cases must hold (note that the combination  $x_{n-1} \in \{0, 1\}$  and  $x_n \in \{\bar{0}, \bar{1}\}$  is impossible):

- (a) Assume  $x_n, x_{n-1} \in \{0, 1\}$ . Then  $a_n = x_n$  and  $b_n = \epsilon$  by (3.1). Consequently,  $x_1 x_2 \cdots x_{n-1}$  is the  $(n-1)$ -length suffix of  $\hat{a}_1 b_1 \hat{a}_2 b_2 \cdots \hat{a}_{n-1} b_{n-1}$  by (3.1).

- (b) Assume  $x_{n-1} \in \{\bar{0}, \bar{1}\}$ . Then  $\bar{a}_n = x_{n-1}$  and  $b_n = x_n$  by (3.1). Consequently,  $x_n x_1 x_2 \cdots x_{n-2}$  is the  $(n-1)$ -length suffix of  $\hat{a}_1 b_1 \hat{a}_2 b_2 \cdots \hat{a}_{n-1} b_{n-1}$  by (3.1). (Because  $b_n$  is the last *unstored* symbol which has been appended to the rear of the queue, it is the last symbol to have been unstored from the front of the queue. Therefore, to restore the queue contents just before  $(a_n, b_n)$  is processed, we delete suffix  $a_n b_n$  from  $x_1 x_2 \cdots x_n$  and prefix the remaining string with  $b_n$ .)

Iterating this reasoning  $n$  times we recover all of  $a_1 a_2 \cdots a_n$ . This proves the lemma.

**Lemma 3.3:** For a word of the form (3.1) with  $|b_1 \cdots b_n| = m$ , we can reconstruct  $a_1 a_2 \cdots a_{m/2}$  from  $b_1 \cdots b_n$ .

**Proof:** Let

$$b_1 b_2 \cdots b_n = x_1 x_2 \cdots x_m, \quad x_i \in \{0, 1, \bar{0}, \bar{1}\} \quad (1 \leq i \leq m).$$

By (3.1) we have  $\hat{a}_1 = x_1$ .

- (a) If  $x_1 \in \{0, 1\}$  then  $a_1 = x_1$  and  $b_1 = \epsilon$ . Consequently,  $x_2 \cdots x_m$  is the  $(m-1)$ -length prefix of  $\hat{a}_2 b_2 \cdots \hat{a}_n b_n$  by (3.1).
- (b) If  $x_1 \in \{\bar{0}, \bar{1}\}$  then  $\bar{a}_1 = x_1$  and  $b_1 = x_2$ . Consequently,  $x_3 \cdots x_m$  is the  $(m-2)$ -length prefix of  $\hat{a}_2 b_2 \cdots \hat{a}_n b_n$  by (3.1).

Iterating this reasoning  $m/2$  times we recover all of  $a_1 a_2 \cdots a_{m/2}$ . This proves the lemma.

**Theorem 3.2.** It requires  $\Omega(n^2)$  time to deterministically simulate one queue by one off-line tape with one-way input.

**Proof.** (I). Assume, by way of contradiction, that an off-line deterministic 1-tape machine  $M$  with one-way input accepts  $L_q$  in time  $T(n) \notin \Omega(n^2)$ . We derive a contradiction by showing that then some incompressible string has too short a description. Without loss of generality, it can be assumed that  $M$  has a semi-infinite storage tape  $[0, \infty)$  on which it writes only 0's and 1's. The positions at time  $t$  are denoted by  $h_1(t)$  and  $h_2(t)$ . By  $t_i$  we denote the time when the  $i$ th input command is polled, i.e.,  $h_1(t_i) = i$  and  $h_1(t_i - 1) = i - 1$ . Fix a constant  $C$  and the word length  $n$  as large as needed to derive the desired contradictions below and such that the formulas in the sequel are meaningful. Below we show that  $T(m) > m^2/C^4$ , for some  $m$ ,  $\sqrt{n}/C \leq m \leq n$ , which contradicts the assumption and proves the theorem.

First, choose an incompressible string  $x \in \{0, 1\}^*$  of length  $n$ . We consider the behavior of  $M$  on a fixed input prefix. This can be any string  $z$  such that  $x = x_1 x_2 \cdots x_n$ ,  $y = y_1 y_2 \cdots y_n$  and  $z = (x_1, y_1)(x_2, y_2) \cdots (x_n, y_n)$ , for some  $y$  such that  $z$  satisfies (3.1). Therefore,  $z \in L_q$ . If many polls occur while the head  $h_2$  is in some small area, then we can show that  $x$  is not incompressible (Case 1). Otherwise, we choose particular  $y_i$ 's, among the possibilities which remain under this constraint, so as to suit the argument in Case 2 below.

*Case 1 (Jammed).* Fix an integer  $m$  such that  $\sqrt{n}/C \leq m \leq n$  (any such integer  $m$  will do) and consider the  $m$ -length prefix  $z(m)$  of  $z$ . By (3.1), if  $z$  is in  $L_q$  then so

is each prefix of  $z$ , so in particular  $z(m) \in L_q$ . Assume, by way of contradiction, that in the accepting computation on  $z(m)$  at least  $2m/C$  polls occur, with  $h_2$  on a particular  $(m/C)$ -length tape segment  $R = [a, a + m/C)$ . Consider the two tape segments  $R_l$  and  $R_r$  of length  $|R|/4$  left and right of  $R$ . Choose positions  $p_l$  in  $R_l$  and  $p_r$  in  $R_r$  with the shortest c.s.'s in their respective tape segments. These c.s.'s must both be shorter than  $m/C^2$ , for if the shortest c.s. in either tape segment is longer than  $m/C^2$  then  $M$  uses  $T(m) \geq m^2/4C^3$  time, which is a contradiction. (If  $0 \leq a < m/4C$  then set  $|R_l| = a$ , so that  $R_l R R_r \subset [0, \infty)$ . Choose  $p_l = 0$  and note that the length of the associated c.s. can be set to 0.) We show that a short program can be constructed which accepts only  $x$ . Let the bits of  $x_1 \cdots x_m$  polled with  $h_2$  outside tape segment  $[p_l, p_r]$ , concatenated in the order in which they occur in  $x$ , form a string  $u$ .

As explained below, we can construct a program to check if a string  $x' \in \{0,1\}^*$  equals  $x$ , using as a description the values of  $n, m, C, a$ , the locations of  $p_l$  and  $p_r$ , the two c.s.'s at  $p_l$  and  $p_r$ , the self-delimiting version of  $u$ , the bits  $x_{m+1} \cdots x_n$ , the final contents of  $[p_l, p_r]$  at time  $t_{m+1}$ , the state of  $M$  at time  $t_{m+1}$  and  $h_2(t_{m+1})$ .

This description of  $x$  requires no more than  $n - \frac{m}{4C}$  bits, for sufficiently large  $C$  and  $n$ . However, this contradicts the incompressibility of  $x$  since  $K(x) \geq n$  and  $m \geq \sqrt{n}/C$ .

To check whether a string  $x'$  equals  $x$ , check  $|x'| = n$  and  $x'_{m+1} \cdots x'_n = x_{m+1} \cdots x_n$ . By the Jamming Lemma (using the above information as related to  $M$ 's processing of the input) reconstruct the contents of  $M$ 's storage tape at time  $t_{m+1}$ , after processing  $z(m) = (x_1, y_1) \cdots (x_m, y_m)$ . Simulate  $M$  from time  $t_{m+1}$  onwards on an input suffix

$$(0, y_{m+1})(0, y_{m+2}) \cdots (0, y_{2m}) \quad (3.2)$$

with  $|y_{m+1}y_{m+2} \cdots y_{2m}| = m$ , and such that  $M$  accepts for the chosen  $y_i$ 's ( $m+1 \leq i \leq 2m$ ). It is easy to see from (3.1), that there is such a suffix (3.2) for which  $M$  accepts if  $x'_1 x'_2 \cdots x'_m = x_1 x_2 \cdots x_m$ . In that case  $x' = x$ , and by (3.1) and Lemma 3.1,  $y_{m+1}y_{m+2} \cdots y_{2m}$  equals the  $m$ -length suffix of  $\hat{x}_1 y_1 \cdots \hat{x}_m y_m$ . By Lemma 3.2, we can retrieve  $x_1 x_2 \cdots x_m$  from this suffix. Suppose, there is a  $x' \neq x$  such that

$$z'(m) = (x'_1, y'_1)(x'_2, y'_2) \cdots (x'_m, y'_m) \quad (3.3)$$

matches the description above, and  $z'(m)$  drives  $M$  into the same configuration at time  $t'_{m+1}$  of  $M$ 's  $(m+1)$ th poll in its computation, as the configuration into which  $z(m)$  drives  $M$  at time  $t_{m+1}$ . Consequently, the concatenation of (3.3) and (3.2) is also accepted by  $M$ . Note, that  $x'$  differs from  $x$  only in the first  $m$  bits, in particular in those bits polled with  $h_2$  positioned in tape segment  $[p_l, p_r]$ . We can cut and paste the computations based on  $z'(m)$  inside  $[p_l, p_r]$  and based on  $z(m)$  outside  $[p_l, p_r]$ , and still have  $M$  accept. The 'cut and paste' computation is accepting up to the  $(m+1)$ th poll because both computations satisfy the description above, and afterwards because the two computations are identical from the  $(m+1)$ st poll onwards. Let the resulting

string composed in the obvious way from  $x_1 \cdots x_m$  and  $x'_1 \cdots x'_m$  be  $\chi(m) = \chi_1 \cdots \chi_m$  with  $\chi_i \in \{x_i, x'_i\}$  ( $1 \leq i \leq m$ ). Above we saw that we can retrieve  $x_1 x_2 \cdots x_m$  from  $y_{m+1} \cdots y_{2m}$ , by Lemma 3.1 and Lemma 3.2. However, this contradicts the acceptance by  $M$  of the cut and paste computation based on  $z(m)$  and  $z'(m)$ , because that entails the retrieval of  $\chi(m) \neq x_1 x_2 \cdots x_m$  from  $y_{m+1} \cdots y_{2m}$  by (3.1), Lemma 3.1 and Lemma 3.2.

*Case 2 (Not jammed).* Let  $n'$  be any integer such that  $\sqrt{n}/C \leq n' \leq n$ . Let  $z(n')$  be the  $n'$ -length prefix of an  $n$ -length input  $z$ . By (3.1), if  $z \in L_q$  then  $z(n') \in L_q$ . Assume, by way of contradiction, that in the accepting computation of  $M$  on  $z(n')$  at most  $2n'/C$  polls occur with  $h_2$  on any particular  $(n'/C)$ -length tape segment  $R = [a, a + n'/C)$ .

We now define the particular input  $z$  we need. Let  $x = x_1 x_2 \cdots x_n$  be as in Case 1. Determine the  $y_i$ 's ( $1 \leq i \leq n$ ) in input  $z = (x_1, y_1) \cdots (x_n, y_n)$  as follows.

- (1) Let  $M$  start its computation with  $y_1 = \epsilon$ . So first  $(x_1, y_1)$  is polled.
- (2) Let  $M$  continue its computation and suppose we have determined  $(x_1, y_1) \cdots (x_{i-1}, y_{i-1})$  and  $M$  polls for the  $i$ th time. Let  $t_i$  be the time at which  $M$  polls  $(x_i, y_i)$ . If  $h_2(t_i) \in [0, n/4)$  then  $y_i = \epsilon$ , else  $y_i \neq \epsilon$ . In the latter case  $y_i$  is determined uniquely from  $(x_1, y_1) \cdots (x_{i-1}, y_{i-1})$  by using the relation  $y_1 y_2 \cdots y_i \leq_{\text{prefix}} \hat{x}_1 y_1 \cdots \hat{x}_{i-1} y_{i-1}$ , that is, using (3.1) and the fact that  $y_1 = \epsilon$  by (1).

We now fix a particular value  $m$  as determined by  $M$ 's computation on  $z$ .

- (a) By contradictory assumption (with  $n' = n$ ), we have that  $\leq n/2$  polls occur on  $[0, n/4)$  and  $\geq n/2$  polls occur on  $[n/2, \infty)$ .
- (b) Since  $T(n) \notin \Omega(n^2)$ , we have  $h_2(t_i) \in [0, n/4)$ , for all  $i$  ( $1 \leq i \leq \sqrt{n}$ ).

Let  $l(t)$  and  $r(t)$  be the number of polls for  $(x_i, y_i)$ 's, with  $h_2(t_i) \in [0, n/4)$  and  $h_2(t_i) \in [n/4, \infty)$  ( $1 \leq t_i \leq t$ ), respectively. By (a) and (b) there is an integer  $m$  such that  $l(t) > r(t)$ , for  $1 \leq t < t_m$ ,  $l(t_m) = r(t_m)$  and  $\sqrt{n}/C \leq m \leq n$ . This  $m$  is the *break even* length where the number of polls left and right of position  $n/4$  on the tape is equal for the first time.

*Claim 1.* As a consequence of this definition of  $m$  and (1) and (2), it follows that  $r(t_m) = |y_1 \cdots y_m| = m/2$  for input prefix

$$z(m) = (x_1, y_1) \cdots (x_m, y_m).$$

Since each prefix of  $z$  satisfies (3.1), we can retrieve  $x_1 \cdots x_{m/4}$  from prefix  $y_1 \cdots y_m$  of  $x_1 y_1 \cdots x_m y_m$  by Lemma 3.3.

*Claim 2.* By definition, all  $y_i$ 's in  $y_1 \cdots y_m$ , which are different from  $\epsilon$ , are polled on  $[n/4, \infty)$ . Since  $l(t_{m/4}) > r(t_{m/4})$ , at most  $m/8$  of the  $x_i$ 's in  $x_1 \cdots x_{m/4}$  are polled on  $[n/4, \infty)$ .

In the computation on the  $m$ -length prefix  $z(m)$  of  $z$ , choose the point  $p$  with the shortest c.s. in  $[n/4 - m/C, n/4)$ . This c.s. is shorter than  $m/C^2$ ; otherwise, the

running time  $T(m) > m^2/C^3$ , which is a contradiction.

As explained below, we can construct a program to check if a string  $x' \in \{0,1\}^*$  equals  $x$ , using as a description the values of  $n$ ,  $m$ , the position of  $p$ , the c.s. at  $p$ , the string  $u$  of concatenated bits of  $x_1 \cdots x_{m/4}$ , polled with  $h_2$  on  $[p, \infty)$  and the string  $x_{(m/4)+1} \cdots x_n$ .

This description of  $x$  requires no more than  $n - \frac{m}{16}$  bits, for sufficiently large  $C'$  and  $n$ . However, this contradicts the incompressibility of  $x$  since  $K(x) \geq n$  and  $m \geq \sqrt{n}/C$ .

To check whether a string  $x'$  equals  $x$ , check  $|x'| = n$  and  $x'_{(m/4)+1} \cdots x'_n = x_{(m/4)+1} \cdots x_n$ . Let  $u'$  be the result of deleting the bits in  $x'$  in the same positions as the ones used to obtain  $u$  from  $x$ . These positions are determined by the crossing sequence at  $p$ . Check  $u' = u$ . If the test is negative then  $x' \neq x$ , else  $x'$  can only differ from  $x$  on positions where  $x_1 \cdots x_{m/4}$ 's bits are polled with  $h_2$  on  $[0, p)$ . Run  $M$  on  $z'(m)$ , that is, the input constructed according to (1), (2), using the  $m$ -length prefix  $x'_1 x'_2 \cdots x'_m$  of a candidate  $x'$ . Whenever  $h_2$  crosses  $p$  we interrupt  $M$  and check if the current  $ID$  in the computation is consistent with the corresponding  $ID$  in the c.s. at  $p$ .

By construction everything matches up to the end of processing input  $z'(m)$ , and  $M$  accepts, if  $x' = x$ . Assume that  $x' \neq x$  matches the description as well. Therefore,  $x'_1 x'_2 \cdots x'_{m/4} \neq x_1 x_2 \cdots x_{m/4}$  and  $x'_i = x_i$  for all  $i$  ( $m/4 + 1 \leq i \leq n$ ). Let the input  $z'(m)$ , based on  $x'_1 x'_2 \cdots x'_m$  and constructed according to (1),(2), be

$$z'(m) = (x'_1, y'_1)(x'_2, y'_2) \cdots (x'_m, y'_m).$$

Let the input based on  $x_1 x_2 \cdots x_m$ , constructed according to (1), (2), be

$$z(m) = (x_1, y_1)(x_2, y_2) \cdots (x_m, y_m).$$

By assumption,  $x'$  and  $x$  differ only on the first  $m/4$  bits, and then only on the bits that are polled left of  $p$ . Let the final accepting position of  $h_2$  for  $M$ 's computation on  $z(m)$  be right of  $p$ . (If it is left of  $p$  interchange  $z$  and  $z'$  below.) Cut and paste the computations on  $z(m)$  and  $z'(m)$  such that  $M$  runs on input  $z'(m)$  with  $h_2$  left of position  $p$ , and  $M$  runs on input  $z(m)$  with  $h_2$  right of position  $p$ . Let  $\zeta(m)$  be the input composed in this way from  $z'(m)$  and  $z(m)$ . By construction, the computation on  $\zeta(m)$  is also an accepting computation of  $M$ . Consequently,  $\zeta(m)$  satisfies (3.1). Then,

$$\zeta(m) = (a_1, b_1)(a_2, b_2) \cdots (a_m, b_m)$$

with  $(a_i, b_i)$  is either  $(x_i, y_i)$  or  $(x'_i, y'_i)$  ( $1 \leq i \leq m$ ). Because both  $z(m)$  and  $z'(m)$  match the description,  $(x'_i, y'_i)$  is polled right of position  $p$  if and only if  $(x_i, y_i)$  is polled right of position  $p$  for all  $i$ ,  $1 \leq i \leq m$ . Therefore,  $y_i = \epsilon$  if and only if  $y'_i = \epsilon$  if and only if  $(x_i, y_i)$  is polled left of position  $p$  if and only if  $(x'_i, y'_i)$  is polled left of position  $p$  for all  $i$ ,  $1 \leq i \leq m$ . Consequently, the sequences of 'unstored' symbols

unequal  $\epsilon$  in the computations on  $z(m)$ ,  $z'(m)$  and  $\zeta(m)$  are equal, that is,

$$b_1 b_2 \cdots b_m = y_1 y_2 \cdots y_m \quad (3.4)$$

By assumption,  $(x_i, y_i)$  and  $(x'_i, y_i)$  are polled left of  $p$  and  $x_i \neq x'_i$  for some  $i$ ,  $1 \leq i \leq m/4$ . Therefore, since  $a_i = x'_i$  if the  $i$ th poll occurs left of  $p$ , and  $a_i = x_i$  if the  $i$ th poll occurs right of  $p$  ( $1 \leq i \leq m/4$ ), in  $M$ 's computation on  $\zeta(m)$ , we have  $a_1 a_2 \cdots a_{m/4} \neq x_1 x_2 \cdots x_{m/4}$ . Because  $x_1 x_2 \cdots x_{m/4}$  is retrieved from  $y_1 \cdots y_m$  by Claim 1, we retrieve  $x_1 x_2 \cdots x_{m/4}$  from  $b_1 b_2 \cdots b_m$  as well, by (3.4), using Lemma 3.3. However, for  $\zeta(m)$  to satisfy (3.1), we have to retrieve  $a_1 a_2 \cdots a_{m/4}$  from  $b_1 b_2 \cdots b_m$ , by virtue of the construction of  $\zeta(m)$  and Claim 1. Consequently,  $b_1 b_2 \cdots b_m$  is not a prefix of  $\hat{a}_1 b_1 \hat{a}_2 b_2 \cdots \hat{a}_m b_m$  as required by (3.1). Hence,  $\zeta(m)$  does not satisfy (3.1), which is a contradiction.

Since  $m \geq \sqrt{n}/C$ , Cases 1 and 2 complete the proof of  $T(n) \in \Omega(n^2)$ .

(II). With the description of  $L_q$  we have already indicated how a queue recognizes this language in real-time.

The theorem follows from (I) and (II).

#### 4. Lower bound for simulating one pushdown by one queue.

In this section, we show that it takes at least  $\Omega(n^{4/3}/\log n)$  time for a one-way input one queue nondeterministic machine to recognize the language  $\{w\#w^R : w \in \{0,1\}^*\}$ .

Because this language can be recognized in linear time by a deterministic pushdown automaton, we can conclude that it takes at least  $\Omega(n^{4/3}/\log n)$  time for a one queue nondeterministic machine to simulate a deterministic pushdown automaton.

The intuition behind the proof is that while the queue machine reads  $w$ , it has to store all the information in some sequential way on the queue. To compare this information with  $w^R$ , the machine will have to go through the queue too many times.

Let  $h_1$  be the read-only head on the one-way input tape. We can view the queue as a tape with two heads  $h_2$  and  $h_3$ . The head  $h_2$  is a read-only, one-way head on the queue. The head  $h_3$  is a write-only, one-way head on the queue. Each time something is put on the queue,  $h_3$  writes and each time something is read from the queue,  $h_2$  reads.

**Theorem 4.1:** A one-way input one queue nondeterministic machine takes time in  $\Omega(n^{4/3}/\log n)$  to accept the language  $\{w\#w^R : w \in \{0,1\}^*\}$ .

**Proof:** Leading to a contradiction, we suppose that there is an algorithm to accept  $L$  in time  $T(n)$  which is not in  $\Omega(n^{4/3}/\log n)$ .

Let  $h_i(t)$  denote the position of head  $i$  at time  $t$  on its respective tape. At time  $t$ , the length of the queue is  $h_3(t) - h_2(t)$ , and the content of the tape between  $h_2(t)$  and  $h_3(t)$  is called the *actual queue*.

Let  $x$  be an incompressible string. We separate  $x$  into blocks:  $x = x_0 x_1 x_2 \cdots x_m$ . Each block  $x_i$  for  $1 \leq i \leq m$  is separated into  $p$  subblocks:  $x_i = x_{i1} x_{i2} x_{i3} \cdots x_{ip}$ . For the proof of the theorem, we take  $m = n^{1/3}$  and  $p = n^{2/3}/k_1 \log n$ , where  $k_1$  is an

appropriately chosen constant. Let  $|x| = n$  and  $|x_0| = n/2$ . Each subblock will have the same length,  $c \log n$ . We look at any computation of the machine on input  $x \# x^R$ .

**Claim 4.2:** If  $n/2 \leq h_1(t) \leq 3n/2$ , then the length of the queue at time  $t$  is at least  $n/2 - \log n$ .

**Proof:** We know that  $K(x_0) \geq n/2 - k_2$  for some constant  $k_2$ . The result follows because  $x_0$  can be described by the content of the queue, the state of the machine and  $h_1(t)$ .

**Claim 4.3:** Let  $t_j$  be the time step when the input head enters the block  $x_j$ . For at least half of the blocks  $x_j$ , we have that  $h_2(t_{j+1}) < h_3(t_j)$ .

**Proof:** Otherwise the algorithm takes time  $\Omega(n^{4/3}/\log n)$ .

The machine needs to remember what it reads on the input and code it in some way on the queue. This notion will be captured by the influence relation defined below. What is written on the queue can be a coding of the input and of the rear of the queue. If  $h_2(t_{j+1}) \leq h_3(t_j)$ , then we have the nice property that a whole block from the input has to be coded sequentially on the queue, since the reading head on the queue doesn't reach where the coding has started. Let's call the blocks which satisfy this last claim the valid blocks.

Now, we define the influence relation. Let  $c_1, c_2, \dots, c_p$  be the cells on the input tape. Let  $d_1, d_2, \dots, d_q$  be the cells on the queue. We say that a cell  $d_j$  is *directly influenced* by a cell  $c_i$  if  $h_1(t) = i$  at the time  $t$  when  $h_3$  writes on  $d_j$ . Similarly, a cell  $d_j$  is *directly influenced* by a cell  $d_i$  if  $h_2(t) = i$  at the time  $t$  when  $h_3$  writes on  $d_j$ .

The *influence relation* is the transitive closure of the direct influence relation. We say that  $c_i$  (or  $d_i$ ) influences  $d_j$  if there is a chain of direct influences from  $c_i$  to  $d_j$ . A block of cells influences a cell if and only if at least one of the cells in the block influences it.

The influence relation will allow us to talk about where the information can be stored on the queue. Notice that during the computation, the content of a cell may still be dependent on some other input cell even if that input cell has no influence on it, due to the finite control of the machine. This minor problem will not cause any trouble.

**Claim 4.4:** For any block  $x_j$  such that  $h_2(t_{j+1}) < h_3(t_j)$ , we have that each cell in  $x_j$  influences a disjoint set of cells on the queue. Moreover, the regions influenced by these cells form an ordered sequence of regions on the actual queue at any later time.

Now we look at what happens when the input head  $h_1$  reads the second part of the input. Let  $t_j'$  denote the time when the head  $h_1$  enters the block  $x_j^R$  corresponding to  $x_j$ .

**Claim 4.5:** There is at least one valid block  $x_j$  such that  $h_2(t_{j-1}') < h_3(t_j')$ . (Remember that  $x_{j-1}^R$  follows  $x_j^R$ .)

**Proof:** Otherwise the algorithm takes time  $\Omega(n^{4/3}/\log n)$ .

In the following two claims, we mention cycles and crossing sequences. A *cycle* is any span of time from time  $t$  to time  $t'$  such that  $h_2(t') = h_3(t)$ . The *crossing sequence* associated with the border between cell  $d_i$  and cell  $d_{i+1}$  is the list of states of the machine when any head goes from cell  $d_i$  to cell  $d_{i+1}$ . Because the tape is in fact a queue, the crossing sequence will have length 2.

Each block influences a series of regions on the tape, one for each cycle of the queue. The *crossing sequence* around a list of regions is the concatenation of the crossing sequences under the border of each region.

**Claim 4.6:** Throughout  $r$  cycles, starting at time  $t_0'$ , the actual queue always has length at least  $n^{2/3} - (r + k_4)\log n$ , for some constant  $k_4$ .

**Proof:** Let  $x_i$  be the block provided by claim 4.5. Let  $x'$  be the string  $x$  where  $x_i$  has been deleted. Because the regions influenced by the  $x_{ij}$  are ordered on the actual queue, and the regions influenced by the  $x_{ij}^R$  are in reverse order, there is only one contiguous region which can be influenced by both a subblock  $x_{ij}$  and its corresponding subblock  $x_{ij}^R$ . We call this region the overlapping region.

At any time after  $t_0'$ , the string  $x$  can be totally described by  $x'$ , the index of  $x_i$ , the actual queue, the crossing sequences around the overlapping region and the content of the regions that were overlapping at each cycle.

**Claim 4.7:** The machine makes  $\Omega(n^{2/3}/\log n)$  cycles after  $t_0'$ .

**Proof:** The string  $x$  can be described by  $x'$ , the index of  $x_i$ , the crossing sequence around the overlapping region and the content of the regions that were overlapping at each cycle. At each cycle, this information is of length  $O(\log n)$ , so it takes  $n^{2/3}/\log n$  cycles to gather enough information. (At the end, we don't need the actual queue, so  $r$  has to be large to compensate.)

By the last two claims, the machine takes time in  $\Omega(n^{4/3}/\log n)$ .

## 5. Simulating more queues by less queues

In this section we study the power of queue machines with different number of queues. We first show that 2 queues are as good as  $k$  queues in the nondeterministic case. This motivates our research concerning a small number of queues. We also show that 1 queue can simulate  $k$  queues in quadratic time, deterministically or nondeterministically. We will provide tight, and almost tight, lower bounds for our simulations mentioned above.

### 5.1. Upper bounds

**Theorem 5.1:** Two pushdown stores can simulate one queue in linear time, both for deterministic and nondeterministic machines.

**Proof:** Let  $P$  be a two pushdown store machine with 2 pushdown stores  $pd\ 1$ ,  $pd\ 2$ . To simulate a queue, every time a symbol is pushed into the queue,  $P$  pushes the same symbol into  $pd\ 1$ . If a symbol is taken from the queue, then  $P$  pops a (the same)

symbol from  $pd\ 2$  if  $pd\ 2$  is not empty. If  $pd\ 2$  is empty then  $P$  first unloads the entire contents of  $pd\ 1$  into  $pd\ 2$  and then pops the top symbol from  $pd\ 2$ . At the end of the input,  $P$  accepts iff the 1 queue machine accepts.

**Theorem 5.2:** Two queues can nondeterministically simulate  $k$  queues for any fixed  $k$  in linear time.

**Proof:** This is actually the same technique Book and Greibach [BG] used to prove the same theorem for tapes. The 2 queue machine guesses the computation of the  $k$ -queue machine computation and put this guess on 1 queue. Then use the other queue to simulate the computation of each of the  $k$  queues of the simulated machine and check its correctness against the guess on the first queue. We refer the reader to [BG] for the details. (This simulation takes  $O(kn)$  time. But it can be improved to *real* time using the methods developed in [BG].)

**Theorem 5.3:** 3 pushdown stores can nondeterministically simulate  $k$  queues in linear time.

**Proof:** Combine ideas from the above 2 theorems. *I.e.*, guess the computation of the  $k$  queue machine, and put the guess into one pushdown store. Save this guess also to another pushdown store (but put a marker on the top). Then simulate a queue and check the correctness of the guess. (The simulation needs 2 pushdowns, one of the pushdowns has the guessed computation saved on the bottom.) After simulating one queue, retrieve the guessed content and again put it into 2 pushdowns. Repeat this process for each queue.

Notice that a strange phenomenon occurs here. When we have 1 queue and 1 pushdown store, 1 queue is better in the sense that 1 queue can accept all the *r.e.* languages but 1 pushdown cannot. However, when we have more pushdown stores, more pushdown stores seems to be better than queues because they are more efficient.

**Theorem 5.4:** One queue can simulate  $k$  queues in quadratic time, both deterministically and nondeterministically.

**Proof:** Here, use some basic simulation schemes.

This also relates to a interesting problem of whether "2 heads (on one tape) are better than 2 tapes (each with one single head)". Vitiányi [Vi3] showed that 2 tapes cannot simulate a queue in real time if one of the tape has only  $o(n)$  cells to use. Our result here shows that 2 pushdowns can simulate a queue in linear time. It would be interesting to know whether this can be done in real time. The question of how to simulate  $k$  deterministic queues by 2 queues (like the Hennie Stearns simulation in the tape case [HS2]) remains open.

## 5.2. Lower bounds

We now prove optimal lower bounds for above simulations. We define the language  $L$ :

$$L = \{ a \& b_0^1 b_1^1 \cdots b_k^1 \# \\ b_0^2 b_0^3 b_1^2 b_2^2 b_1^3 b_3^2 \cdots b_{2i}^2 b_i^3 b_{2i+1}^2 \cdots b_{k-1}^2 b_{(k-1)/2}^3 b_k^2 \}$$

$$b_0^4 b_{(k+1)/2}^3 b_1^4 b_2^4 b_{(k+3)/2}^3 b_3^4 \cdots b_{2i \bmod (k+1)}^4 b_i^3 b_{(2i+1) \bmod (k+1)}^4 \cdots b_{k-1}^4 b_k^3 b_k^4$$

$n \mathcal{S} a$

$$\{ b_i^1 = b_i^2 = b_i^3 = b_i^4 \text{ for } i = 0, \dots, k \text{ for any odd } k \}$$

The length of each  $b_i^j$  is a fixed constant  $C$ . The superscripts of  $b_i$ 's are used only to facilitate later discussions.  $L$  can be considered as a modified version of a language used in [Ma]. We have added a string  $a$  on the both ends. The purpose of  $a$  is to prevent the queue from shrinking since if we choose  $a$  to be a long random string then before the second  $a$  is read the size of the queue has to be larger than  $|a|$ . We have to prevent the queue from shrinking because otherwise the crossing sequence argument would not work. In order to prove the lower bounds for queues new techniques, in addition to those used in [Ma, LV], are required.

**Theorem 5.5:** Simulating two deterministic queues by one nondeterministic queue requires  $\Omega(n^2/\log^2 n \log \log n)$  time.

**Proof:** We will show that  $L$  defined above requires  $\Omega(n^2/\log^2 n \log \log n)$  time on a one-queue nondeterministic machine (always with an extra 1-way input tape). Since  $L$  can obviously be accepted by a two-queue deterministic machine in linear time, the theorem will follow.

Now, aiming at a contradiction, assume that a one-queue machine  $M$  accepts  $L$  in  $o(n^2/\log^2 n \log \log n)$  time. Only for the notational convenience, we think the queue of  $M$  as a *circular tape* with just *one queue head*, which combines the push head and the pop head. The head moves clockwise uniformly. The circular tape can augment (insert a tape square) or shrink (delete a tape square) at constant cost in order to mimic a queue. We call it *Queue* and write  $|Queue(t)|$  to denote its length at time  $t$ . Name the input head  $h_1$  and the queue head  $h_2$ . Initially the *Queue* is a point, a degenerated ring.

Choose a large  $n$  and a  $C > 10 |M| + 10$  so that all the subsequent formulas make sense. Choose an incompressible string  $X \in \{0,1\}^{2^n}$ . Let  $X = X' X''$  where  $|X'| = |X''|$ . Equally divide  $X''$  into  $k+1 = n/C \log \log n$  parts,  $X'' = x_0 x_1 \cdots x_k$ , each  $C \log \log n$  long. Consider a word  $w \in L$  where  $a = X'$  and  $b_i^j = x_i$  for  $1 \leq j \leq 5$  and  $0 \leq i \leq k$ . Fix a shortest accepting path,  $P$ , of  $M$  on  $w$ .

Consider only the path  $P$ . Let  $t_{\mathcal{S}}$  be the time when  $h_1$  reaches the first  $\mathcal{S}$ ,  $t'_{\mathcal{S}}$  be the time  $h_1$  reaches the second  $\mathcal{S}$ , and  $t_{\#}$  be the time when  $h_1$  reaches  $\#$ .

**Claim 1.**  $|Queue(t_{\mathcal{S}})| \geq n/2$ .

**Proof.** If not, we can conclude that  $K(X) < |X|$  as follows. For every  $Y$  such that  $|Y| = |X|$ , let  $Y = a' y_0 \cdots y_k$ . Replace the last  $a$  after the  $\mathcal{S}$  sign in  $w$  by  $a'$ . Using the (short) description of the queue, start to simulate  $M$  from time  $t_{\mathcal{S}}$ . By a standard argument,  $Y = X$  iff  $M$  accepts. Therefore  $K(X) < |X|$ , a contradiction.

By a similar argument as in Claim 1, we derive Claim 2.

*Claim 2.*  $|Queue(t)| \geq n/2$  for every  $t_{\theta} \leq t \leq t'_{\theta}$ .

*Claim 3.* The crossing sequence from time  $t_{\theta}$  to  $t'_{\theta}$  is shorter than  $\frac{n}{C^{10} \log^2 n \log \log n}$ .

*Proof.* Follows directly from Claim 2.

If the *Queue* cells which are scanned or created by  $h_2$ , while  $h_1$  is scanning  $b_i^j$ , are in  $Q = \{q_1, \dots, q_u\}$ , then we say that  $b_i^j$  is *mapped into*  $Q$ . Notice that  $b_i^j$  is at first mapped into a set,  $Q$ , of consecutive sequence of cells. But, different from a regular Turing machine tape,  $Q$  may become disconnected because other queue squares can be inserted later. We say that  $b_i^j$  is *sequentially mapped* if, while  $h_1$  scans  $b_i^j$ ,  $h_2$  did not scan any *Queue* cell twice (leave and re-enter), that is,  $h_2$  did not make a round trip on *Queue*. We say that  $b_i^j$  is *majorly mapped into*  $Q$  if  $b_i^j$  is sequentially mapped, and there are two substrings,  $u$  and  $v$ , of  $b_i^j$  which are mapped into  $Q$  and  $|u| + |v| \geq \frac{|b_i^j|}{2}$ .

*Remark.* According to above definition, a  $b_i^j$  can be majorly mapped into two disjoint sets.

*Claim 4.* At time  $t_{\theta}$ , *Queue* can be cut into two segments,  $S_1$  and  $S_2$ , such that

- (1)  $S_1 \cap S_2 = \phi$  and  $S_1 \cup S_2 = Queue$ ;
- (2)  $k/4$   $b_i^1$ 's, say  $b_{i_1}^1, \dots, b_{i_{k/4}}^1$ , are majorly mapped into  $S_1$ , and  $k/4$   $b_i^1$ 's, say  $b_{j_1}^1, \dots, b_{j_{k/4}}^1$ , are majorly mapped into  $S_2$ .  
 $\{b_{i_1}^1, \dots, b_{i_{k/4}}^1\} \cap \{b_{j_1}^1, \dots, b_{j_{k/4}}^1\} = \phi$ .
- (3)  $|S_1|, |S_2| \geq n/C^2$ ;

*Proof.* In our proof only properties (1) and (2) are used, (3) is stated for the sake of completeness. We will only give proofs of (1) and (2). The proof of (3) is very similar to the last part of this proof and we leave the proof of (3) to the interested readers.

First we show that  $\frac{99k}{100} b_i^1$  are sequentially mapped. By Claim 2, for  $t_{\theta} \leq t \leq t'_{\theta}$ , we always have  $|Queue(t)| \geq n/2$ . Therefore, if after time  $t_{\theta}$  more than  $k/100$   $b_i^1$ 's are not sequentially mapped, then on each of them  $M$  must spend at least  $n/2$  time to go around the *Queue*. Altogether  $M$  would be spending  $\Omega(n^2/\log \log n)$  time, a contradiction. Hence at least  $99k/100$   $b_i^1$ 's are sequentially mapped.

Now we can easily choose two points  $p, q$  on *Queue* to cut *Queue* into two parts  $S_1$  and  $S_2$  such that (1) and (2) in the claim are true.

>From now on, we will always consider the *Queue* to be partitioned as  $S_1$  and  $S_2$ . The sizes of  $S_1$  and  $S_2$  may increase or decrease. If anything is inserted in the intersection point of  $S_1$  and  $S_2$  then it does not matter in which set we place the inserted *Queue* cell.

The next claim is a simple generalization of a theorem proved by Maass in [M] (Theorem 3.1). The proof of this claim is a simple reworking of the proof in [M].

*Claim 5.* Let  $S$  be a sequence of numbers from  $\{0, \dots, k-1\}$ , where  $k=2^l$  for some  $l$ . Assume that every number  $b \in \{0, \dots, k-1\}$  is somewhere in  $S$  adjacent to the number  $2b \pmod{k}$  and  $2b \pmod{k} + 1$ . Then for every partition of  $\{0, \dots, k-1\}$  into two sets  $G$  and  $R$  such that  $S = G \cup R$  and  $|G|, |R| > k/4$  there are at least  $k/(c \log k)$  (for some fixed  $c$ ) elements of  $G$  that occur somewhere in  $S$  adjacent to a number from  $R$ .

A  $k/\sqrt{\log k}$  upper bound corresponding to the lower bound in this lemma is contained in [Li]. A more general, but weaker, upper bound can be found in [KI]. Notice that any sequence  $S$  in  $L$  satisfies the requirements in Claim 5.

*Claim 6.* At time  $t_{\#}$ , the  $b_i$ 's between  $\#$  and the second  $\mathcal{E}$  are mapped into *Queue* in the following way: either

- (1) a set,  $\bar{S}_1$ , of  $k/c \log k$   $b_j$ 's, which belong to  $\{b_{j_1}^1, \dots, b_{j_{k/4}}^1\}$ , are mapped into  $S_1$ ;  
or
- (2) a set,  $\bar{S}_2$ , of  $k/c \log k$   $b_i$ 's, which belong to  $\{b_{i_1}^1, \dots, b_{i_{k/4}}^1\}$ , are mapped into  $S_2$ .

Where  $c \ll C$  is a small constant as used in Claim 5.

*Proof.* By Claim 3 we can assume that from time  $t_{\#}$  to time  $t_{\mathcal{E}}$ ,  $h_2$  made less than  $\frac{k}{C^2 \log k}$  round trips on *Queue*. Therefore by the nature of the queue, only  $\frac{2k}{C^2 \log k}$   $b_i^j$ 's can be mapped into both  $S_1$  and  $S_2$ . Also since  $h_2$  can alternate between  $S_1$  and  $S_2$  less than  $\frac{2k}{C^2 \log k}$  times, we complete the proof by applying Claim 5.

Without loss of generality, assume that (1) of Claim 6 is true.

*Claim 7.* Let  $t_{end}$  be the time  $M$  ends. Either

- (a) There exists a time  $t_{\mathcal{E}} \leq t_0 \leq t_{end}$  such that  $|Queue(t_0)| \leq \frac{n}{C^{10} \log n}$  and the crossing sequence from  $t_{\mathcal{E}}$  to  $t_0$  is shorter than  $\frac{n}{C^{10} \log n \log \log n}$ ; or
- (b)  $>$  From time  $t_{\mathcal{E}}$  to time  $t_{end}$  the length of the crossing sequence is shorter than  $\frac{n}{C^{10} \log n \log \log n}$ .

*Proof.* If (a) and (b) are both false, then  $M$  spends  $\Omega\left(\frac{n^2}{\log^2 n \log \log n}\right)$  time, a contradiction.

By Claim 3 the crossing sequence is shorter than  $\frac{n}{C^{10} \log^2 n \log \log n}$  before time  $t_{\mathcal{E}}$ . Record this crossing sequence. For every  $j, k$ , if  $b_j^k \in \bar{S}_1$ , then  $b_j^1$  is majorly mapped into  $S_2$ . Let  $u_j, v_j$  be the substrings of  $b_j^1$  such that

- (i)  $|u_j| + |v_j| \geq \frac{|b_j^1|}{2} - 1$  and
- (ii)  $u_j, v_j$  are mapped into  $S_2$ .

Let  $S_{uv} = \{u_j, v_j \mid b_j^k \in \bar{S}_1 \text{ for some } k > 1\}$ . Notice that  $\sum_{u_j, v_j \in S_{uv}} (|u_j| + |v_j|) \geq \frac{n}{3C \log n}$ .

Now we describe a program which reconstructs  $X$  with less than  $|X|$  information. Consider every  $Y$  such that  $|Y| = |X|$  and  $Y = a y_0 \cdots y_k$  for some  $y_0 \cdots y_k$ .

- (1) Check if  $Y$  is the same as  $X$  at positions other than those places occupied by  $u_i, v_i \in S_{uv}$ .
- (2) If (1) is true, then construct the input,  $w_Y$ , as before except with  $x_i$  replaced by  $y_i$  for  $i = 0, 1, \dots, k$ .
- (3) Run  $M$  following path  $P$  up to time  $t_g$ .
- (4) We distinguish between two cases according to Claim 7.

*Case 1.* (b) of Claim 7 is true. Record the crossing sequence from  $t_g$  to  $t_{end}$ . Continue to run  $M$  from  $t_g$  to  $t_{end}$  such that  $h_2$  never goes into  $S_2$ . Whenever  $h_2$  reaches the border of  $S_2$  it matches the current  $ID$  with the crossing sequence. If they match  $M$  jumps over  $S_2$  and, using the next  $ID$  on the other side of  $S_2$  to start from,  $M$  continues until time  $t_{end}$ .

*Case 2.* (a) of Claim 7 is true. Record the crossing sequence from time  $t_g$  to time  $t_0$  and the contents of  $S_2$  at time  $t_0$ . Simulate  $M$ , with  $h_2$  staying outside of  $S_2$ , from time  $t_g$  to time  $t_0$  similar to Case 1. At time  $t_0$ ,  $M$  puts the (short) contents of  $S_2$  in the position of  $S_2$  and then finishes the computation in the normal way.

- (5) By the end  $M$  accepts iff  $Y = X$ . Notice that since  $M$  is nondeterministic, by "accept" we mean that there is an accepting path.

Now the information we used in this program is only the following.

- (1)  $X - S_{uv}$ , plus the information to describe the relative locations of  $b_j \in \bar{S}_1$  in  $X$  and the relative locations of  $u_j, v_j$  in  $b_j \in \bar{S}_1$ . Using the coding method described in the previous part of the paper, this would require at most

$$|X| - \frac{n}{2c \log n} + \frac{n}{\frac{C}{3} c \log n} + o\left(\frac{n}{\log n}\right),$$

where the second term is for the  $u_i, v_i$  in  $S_{uv}$ , the third term is for the information to describe the relative positions of  $b_j \in \bar{S}_1$ , and the last term is for the information needed to describe the relative positions of  $u_i, v_i$  in each  $b_i$

- (2) Description of the crossing sequence, of length less than  $\frac{n}{C^9 \log n \log \log n}$ , around  $S_2$ . Again by the method used in previous part of this paper, this requires at most

$$\frac{n}{C^8 \log n \log \log n} \text{ bits.}$$

(3) Description of the contents of  $S_1$  at time  $t_0$  when (a) of Claim 6 is true. But  $|Queue(t_0)| \leq \frac{n}{C^{10} \log n}$ .

(4) Extra  $O(\log n)$  bits to describe the program discussed above.

The total is less than  $|X|$ . Therefore  $K(X) < |X|$ , a contradiction.

*Corollary.* Simulating two deterministic tapes by one nondeterministic queue requires  $\Omega(n^2/\log^2 n \log \log n)$ .

*Proof.* Since  $L$  can also be accepted by a two tape Turing machine in linear time.

**Theorem 5.6:** It requires  $\Omega(n^2)$  time to simulate two deterministic queues by one deterministic queue.

*Proof Idea.* Define a language  $L_1$  as follows. (Below,  $a, x_i, y_i \in \{0,1\}^*$ .)

$$L_1 = \{ a \& x_1\$x_2\$ \cdots \$x_k\#y_1\$ \cdots \$y_l\# (1^{i_1}, 1^{j_1})(1^{i_2}, 1^{j_2}) \cdots (1^{i_s}, 1^{j_s}) \& a \mid \\ x_p = y_q \& (p = i_1 + \dots + i_t, q = j_1 + \dots + j_t) \& 1 \leq t \leq s \}.$$

$L_1$  can be accepted by a two queue deterministic machine in linear time. But using the techniques in Theorem 1 and in the proof of one deterministic Turing machine tape requiring square time for this language (See [LV]), it can be shown that  $L_1$  requires  $\Omega(n^2)$  for a one queue deterministic machine. We omit the proof.

*Remark.* The above lower bounds are similar to the case of one tape vs two tapes [Ma, LV]. However, the proofs require special techniques to handle the queues. Still we do not have a lower bound as good as in the nondeterministic tape case [LV, GKS]. We feel that some improvement should be possible.

## References

- [Aa] Aanderaa, S.O., "On k-tape versus (k-1)-tape real-time computation," in *Complexity of Computation*, ed. R.M. Karp, SIAM-AMS Proceedings, vol. 7, pp. 75-96, American Math. Society, Providence, R.I., 1974.
- [BGW] Book, R., S. Greibach, and B. Wegbreit, "Time- and tape-bound Turing acceptors and AFL's," *J. Computer and System Sciences*, vol. 4, pp. 606-621, 1970.
- [Ch] Chaitin, G.J., "Algorithmic Information Theory," *IBM J. Res. Dev.*, vol. 21, pp. 350-359, 1977.
- [DGPR] Duris, P., Z. Galil, W. Paul, and R. Reischuk, "Two nonlinear lower bounds for on-line computations," *Information and Control*, vol. 60, pp. 1-11, 1984.
- [GKS] Galil, Z., R. Kanhan, E. Szemerédi, "On nontrivial separators for k-page graphs and simulations by nondeterministic one-tape Turing machines," in *Proceedings 18th Annual ACM Symposium on Theory of Computing*, 1986.

- [HS1] Hartmanis, J. and R.E. Stearns, "On the computational complexity of algorithms," *Trans. Amer. Math. Soc.*, vol. 117, pp. 285-306, 1969.
- [HS2] Hennie, F.C. and R.E. Stearns, "Two tape simulation of multitape Turing machines," *J. Ass. Comp. Mach.*, vol. 4, pp. 533-546, 1966.
- [HU] Hopcroft, J.E. and J.D. Ullman, *Formal Languages and their Relations to Automata*, Addison-Wesley, 1969.
- [Kl] Klawe, M., "Limitations on explicit construction of expanding graphs," *SIAM J. Comp.*, vol. 13, no. 4, pp. 156-166, 1984.
- [Ko] Kolmogorov, A.N., "Three approaches to the quantitative definition of information," *Problems in Information Transmission*, vol. 1, no. 1, pp. 1-7, 1965.
- [Li1] Li, M., "Simulating two pushdowns by one tape in  $O(n^{1.5} (\log n)^{0.5})$  time," 26th Annual IEEE Symposium on the Foundations of Computer Science, 1985.
- [Li2] Li, M., "Lower Bounds in Computational Complexity," Ph.D. Thesis, Report TR-85-663, Computer Science Department, Cornell University, march 1985.
- [Li3] Li, M., "Lower bounds by Kolmogorov-complexity", 12th ICALP, Lecture Notes in Computer Science, 194, pp. 383-393, 1985.
- [LV] Li, M. and P.M.B. Vitányi, "Tape versus queue and stacks: The lower bounds," Submitted for publication.
- [LS] Leong, B.L. and J.I. Seiferas, "New real-time simulations of multi-head tape units," *J. Ass. Comp. Mach.*, vol. 28, pp. 166-180, 1981.
- [Ma] Maass, W., "Combinatorial lower bound arguments for deterministic and nondeterministic Turing machines," *Trans. Amer. Math. Soc.*, 292,2, pp. 675-693, 1985. (Preliminary Version "Quadratic lower bounds for deterministic and nondeterministic one-tape Turing machines," pp 401-408 in Proceedings 16th ACM Symposium on Theory of Computing, 1984.)
- [PSS] Paul, W.J., J.I. Seiferas, and J. Simon, "An information theoretic approach to time bounds for on-line computation," *J. Computer and System Sciences*, vol. 23, pp. 108-126, 1981.
- [Pa] Paul, W.J., "On-line simulation of  $k+1$  tapes by  $k$  tapes requires nonlinear time," *Information and Control*, pp. 1-8, 1982.
- [So] Solomonov, R., *Information and Control*, vol. 7, pp. 1-22, 1964.
- [Vi1] Vitányi, P.M.B., "One queue or two pushdown stores take square time on a one-head tape unit," Computer Science Technical Report CS-R8406, CWI, Amsterdam, March 1984.
- [Vi2] Vitányi, P.M.B., "An  $N^{1.618}$  lower bound on the time to simulate one queue or two pushdown stores by one tape," *Information Processing Letters*, vol. 21, pp. 147-152, 1985.
- [Vi3] Vitányi, P.M.B., "On two-tape real-time computation and queues," *J. Computer and System Sciences*, vol. 29, pp. 303-311, 1984.

OFFICIAL DISTRIBUTION LIST

1985

Director 2 Copies  
Information Processing Techniques Office  
Defense Advanced Research Projects Agency  
1400 Wilson Boulevard  
Arlington, VA 22209

Office of Naval Research 2 Copies  
800 North Quincy Street  
Arlington, VA 22217  
Attn: Dr. R. Grafton, Code 433

Director, Code 2627 6 Copies  
Naval Research Laboratory  
Washington, DC 20375

Defense Technical Information Center 12 Copies  
Cameron Station  
Alexandria, VA 22314

National Science Foundation 2 Copies  
Office of Computing Activities  
1800 G. Street, N.W.  
Washington, DC 20550  
Attn: Program Director

Dr. E.B. Royce, Code 38 1 Copy  
Head, Research Department  
Naval Weapons Center  
China Lake, CA 93555

Dr. G. Hopper, USNR 1 Copy  
NAVDAC-OOH  
Department of the Navy  
Washington, DC 20374

E N D

10-86

D T / C