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# AN EFFICIENT ALGORITHM FOR MAXDOMINANCE, WITH APPLICATIONS

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**Abstract.** Given a planar set  $S$  of  $n$  points, *maxdominance* problems consist of computing, for every  $p \in S$ , some function of the maxima of the subset of  $S$  that is dominated by  $p$ . A number of geometric and graph-theoretic problems can be formulated as maxdominance problems, including the problem of computing a minimum independent dominating set in a permutation graph, the related problem of finding the shortest maximal increasing subsequence, the problem of computing a maximum independent set in an overlap (and hence circle) graph, the problem of enumerating restricted empty rectangles, and the related problem of computing the largest empty rectangle. We give an algorithm for optimally solving a class of maxdominance problems. A straightforward application of our algorithm yields improved time bounds for the above-mentioned problems. The techniques used in the algorithm are of independent interest, and include a linear-time tree computation that is likely to arise in other contexts.

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## 1. Introduction

A point  $p$  is said to *dominate* a point  $q$  iff  $X(p) \geq X(q)$ ,  $Y(p) \geq Y(q)$ , and  $p \neq q$ , where  $X(p)$  and  $Y(p)$  respectively denote the  $x$  and  $y$  coordinates of point  $p$ . If  $S$  is a set of points and  $p$  is a point, we use  $DOM_S(p)$  to denote the subset of points in  $S$  that are dominated by point  $p$ . A point of  $S$  is a *maximum in  $S$*  iff no other point of  $S$  dominates it. We use  $MAX(S)$  to denote the set of maxima of  $S$ , listed by increasing  $x$  coordinates (and hence by decreasing  $y$  coordinates). We abbreviate  $MAX(DOM_S(p))$  as  $MD_S(p)$ . A number of geometric and graph-theoretic problems can be formulated as one of the following two *maxdominance* problems **P1** and **P2** (problem **P2** being substantially more difficult than **P1**).

**Problem P1.** Given a set  $S$  of  $n$  points in the plane, compute  $MD_S(p)$  for every  $p \in S$ .

We solve the above problem in  $O(n \log n + t)$  time where  $t$  is the size of the output, i.e.

$$t = \sum_{p \in S} |MD_S(p)|.$$

**Problem P2.** For a set  $S$  of points in the plane with a real weight  $w(p)$  associated with every  $p \in S$ , the problem is to compute the *label* and *predecessor* of every point in  $S$ , where the *label* function is defined as follows:

$$label(p) = w(p) \quad \text{if } DOM_S(p) = \emptyset,$$

$$label(p) = w(p) + \text{Min}\{label(q) : q \in MD_S(p)\} \quad \text{otherwise.}$$

The *predecessor* of point  $p$  is any one of the points which gave  $p$  its label, i.e. it is a point  $q \in MD_S(p)$  such that  $label(p) = w(p) + label(q)$  (if  $DOM_S(p) = \emptyset$  then  $p$  has no predecessor).

We solve problem **P2** in  $O(n \log n)$  time and  $O(n)$  space, which is optimal since sorting is a trivial special case of **P2**.

It is the algorithm for **P2** that is the main contribution of this paper (**P1** is solved by a much simplified version of the algorithm for **P2**).

The paper is organized as follows. Section 2 establishes some preliminary results, and Section 3 gives a result on tree computations which is needed in our solution to **P2** (it is also of

independent interest). Section 4 gives our  $O(n \log n)$  time,  $O(n)$  space algorithm for problem P2. Section 5 gives an  $O(n \log n + t)$  algorithm for problem P1. Section 6 lists problems for which improved complexity bounds follow from our results, and Section 7 concludes.

## 2. Preliminaries

Throughout this section,  $L$  and  $R$  are two planar sets of points separated by a vertical line and such that  $L$  is to the left of  $R$ ;  $S$  denotes  $L \cup R$ . To simplify the exposition, we assume that no two points have same  $x$  coordinate (similarly for  $y$  coordinates).

Recall that in the list  $MD_S(p)$ , the points are in increasing  $x$  coordinate value. For every  $p \in S$ ,  $leader_S(p)$  denotes the leftmost (i.e. highest) point in  $MD_S(p)$  (if  $MD_S(p) = \emptyset$  then  $leader_S(p) = \emptyset$ ). In Figure 1,  $MD_R(p) = \{u, v, w\}$ ,  $MD_S(p) = \{b, e, d, c, u, v, w\}$ ,  $leader_R(p) = u$ , and  $leader_S(p) = b$ .

For every  $p \in R$ ,  $Strip_L(p, R)$  denotes the points of  $L$  that are below  $p$  and above  $leader_R(p)$ ;  $Begin_L(p, R)$  and  $End_L(p, R)$  denote the leftmost (i.e. highest) and rightmost (i.e. lowest) points on  $MAX(Strip_L(p, R))$ , respectively (if  $Strip_L(p, R) = \emptyset$  then  $Begin_L(p, R) = End_L(p, R) = \emptyset$ ). For example, in Figure 1,  $Strip_L(p, R) = \{a, b, c, d, e, f\}$ ,  $MAX(Strip_L(p, R)) = \{b, e, d, c\}$ ,  $Begin_L(p, R) = b$ , and  $End_L(p, R) = c$ . Observe that for every  $p \in R$ , the list  $MD_S(p)$  is the concatenation of  $MAX(Strip_L(p, R))$  with  $MD_R(p)$ .

We define  $G(S)$  as the directed acyclic graph whose vertex set is  $S$  and such that  $(p, q)$  is an edge in  $G(S)$  iff there exists a point  $w \in S$  such that  $q$  immediately follows  $p$  on the list  $MD_S(w)$ , in which case we say that edge  $(p, q)$  is *caused by*  $w$ . An edge may be caused by more than one point, but  $G(S)$  has a single copy of such an edge. In Figure 1, edge  $(u, v)$  is in  $G(S)$  and is caused by points  $k, p, g$  and  $h$ . Note also that  $(u, w)$  is not an edge of  $G(S)$ . Let  $E(L, R)$  be the subset of edges of  $G(S)$  that have both ends in  $L$  and are caused by at least one point in  $R$ . That is,

$E(L,R) = \{(p,q) : p \in L, q \in L, (p,q) \text{ is caused by some } w \in R\}$ .

**Observation 1.** The graph  $(L, E(L,R))$  is a forest.

**Proof.** A node in this graph has out-degree at most one.  $\square$

Note that for every  $p \in R$ ,  $MAX(Strip_L(p,R))$  is the path in the forest  $(L, E(L,R))$  from  $Begin_L(p,R)$  to  $End_L(p,R)$ .

Let  $CROSS(L,R)$  be the subset of edges of  $G(S)$  that have one endpoint in  $L$  and one in  $R$ .

**Observation 2.**  $|CROSS(L,R)| \leq |R|$ .

**Proof.** An edge in  $CROSS(L,R)$  can only be caused by a point in  $R$ . Moreover, a point in  $R$  can cause at most one edge in  $CROSS(L,R)$ . Thus  $|CROSS(L,R)| \leq |R|$ .  $\square$

Note that if  $p \in R$  causes the edge  $(c,u) \in CROSS(L,R)$ , then  $c = End_L(p,R)$  (see Figure 1).

Two points  $p$  and  $q$  are *comparable* iff one of them dominates the other. A set of points forms a *chain* iff every two points in it are comparable.  $MAXREV(S)$  denotes the subset of  $S$  such that  $p \in MAXREV(S)$  iff no other point of  $S$  is both above  $p$  and to its left. We assume that the elements of  $MAXREV(S)$  are listed by increasing  $x$  coordinates (and hence by increasing  $y$  coordinates, since they form a chain). In Figure 1,  $MAXREV(R) = \{l, u, k\}$ .

**Lemma 1.** Given the lists  $Q_L$  and  $Q_R$  containing the points of  $L$  and  $R$ , respectively, sorted by increasing  $y$  coordinates,  $E(L,R)$  and  $CROSS(L,R)$  can be computed in  $O(|L| + |R|)$  time. In addition, for all  $p \in R$ ,  $Begin_L(p,R)$  and  $End_L(p,R)$  can also be computed in  $O(|L| + |R|)$  time.

**Proof.** Let  $Q_R = (q_1, \dots, q_{|R|})$ ,  $Y(q_1) < \dots < Y(q_{|R|})$ . Initialize  $E(L,R)$  and  $CROSS(L,R)$  to  $\emptyset$ . We compute the edges in  $E(L,R)$  by scanning the list  $Q_R$ , maintaining on a stack  $STACK$  the  $MAXREV$  of the subset of  $R$  encountered so far by the scan; i.e. when we are at  $q_i$ ,  $STACK$  contains the elements of  $MAXREV(\{q_1, \dots, q_i\})$  stored by increasing  $y$  coordinates. Note that  $q_i$  is the highest point in  $\{q_1, \dots, q_i\}$  and hence it belongs to  $MAXREV(\{q_1, \dots, q_i\})$  and is at the top of  $STACK$ . When the scan advances from  $q_i$  to  $q_{i+1}$ , we do the following: we add to  $E(L,R)$  and  $CROSS(L,R)$  the edges that are caused by  $q_{i+1}$  and are not caused by any of  $\{q_1, \dots, q_i\}$  (i.e. the

"new" edges), update the contents of *STACK* so that it contains  $MAXREV(\{q_1, \dots, q_{i+1}\})$ , and compute  $Begin_L(q_{i+1}, R)$  and  $End_L(q_{i+1}, R)$ . The details are as follows.

- (1) Obtain the elements of  $Strip_L(q_1, R)$  in sorted order. This takes  $O(|Strip_L(q_1, R)|)$  time by scanning  $Q_L$  until a point of  $L$  higher than  $q_1$  is reached. (Note. Since  $leader_R(q_1) = \emptyset$ ,  $Strip_L(q_1, R) = DOM_L(q_1)$ .) Compute  $MAX(Strip_L(q_1, R))$ ; since the points in  $Strip_L(q_1, R)$  are already sorted, this takes  $O(|Strip_L(q_1, R)|)$  time [OV]. Add  $|MAX(Strip_L(q_1, R))| - 1$  edges to  $E(L, R)$ , one for each pair of adjacent points in  $MAX(Strip_L(q_1, R))$ ; i.e. if  $q$  immediately follows  $p$  in  $MAX(Strip_L(q_1, R))$  then we add edge  $(p, q)$  to  $E(L, R)$ . If  $Strip_L(q_1, R) \neq \emptyset$  then set  $Begin_L(q_1, R)$  and  $End_L(q_1, R)$  to be the leftmost and rightmost points on  $MAX(Strip_L(q_1, R))$ , respectively. If  $Strip_L(q_1, R) = \emptyset$  then set  $Begin_L(q_1, R)$  and  $End_L(q_1, R)$  to be  $\emptyset$ .

Set  $i=1$  and repeat the following Steps (2)-(5) until  $i > |R|$ :

- (2) Advance along  $Q_L$  until a point of  $L$  higher than  $q_{i+1}$  is reached. The sequence of points encountered, excluding the last one, yields the subset  $H$  of points in  $L$  that are above  $q_i$  and below  $q_{i+1}$ , sorted by their  $y$  components. Compute  $MAX(H)$ ; since the points in  $H$  are already sorted, this takes  $O(|H|)$  time [OV]. Add to  $E(L, R)$  an edge for each consecutive pair of points in the list  $MAX(H)$  (these edges of  $E(L, R)$  are caused by  $q_{i+1}$  and not caused by any of  $\{q_1, \dots, q_i\}$ ). If  $q_{i+1}$  dominates  $q_i$  then go to Step (3), otherwise go to Step (4).
- (3) Since  $q_{i+1}$  dominates  $q_i$ ,  $q_i$  is  $leader_R(q_{i+1})$  and therefore  $H = Strip_L(q_{i+1}, R)$  and all the new edges of  $E(L, R)$  caused by  $q_{i+1}$  were already added in Step (2). If  $H \neq \emptyset$  then set  $Begin_L(q_{i+1}, R)$  (resp.  $End_L(q_{i+1}, R)$ ) to be the leftmost (resp. rightmost) point on  $MAX(H)$ , then add to  $CROSS(L, R)$  the edge  $(End_L(q_{i+1}, R), q_i)$ . If  $H = \emptyset$  then set  $Begin_L(q_{i+1}, R) = End_L(q_{i+1}, R) = \emptyset$ . Go to Step (5).
- (4) Since  $q_{i+1}$  does not dominate  $q_i$ ,  $q_{i+1}$  is above and to the left of  $q_i$ : Pop from *STACK* all the points that are below and to the right of  $q_{i+1}$ , and let  $\beta_1, \dots, \beta_k$  be the sequence of points so popped (see Figure 2). Note that  $\beta_1 = q_i$ , and that the  $\beta_j$ 's form a chain and are the top  $k$

points on  $MAXREV(\{q_1, \dots, q_i\})$ . Let  $U_0$  denote  $MAX(H)$ , and let  $U_j$  denote  $MAX(Strip_L(\beta_j, R))$  ( $1 \leq j \leq k$ ). Sub-step (4.1) below computes  $Begin_L(q_{i+1}, R)$  and  $End_L(q_{i+1}, R)$ , while sub-step (4.2) finds any additional edges of  $E(L, R)$  that are caused by  $q_{i+1}$  (for example, an edge between the rightmost point of  $U_j$  and the point immediately to its right on  $U_{j+1}$ ). We do not need to add to  $CROSS(L, R)$  the edge (if there is one) caused by  $q_{i+1}$ , because such an edge would also be caused by  $\beta_k$  and thus would already have been added when processing  $\beta_k$ .

(4.1) If  $\bigcup_{j=0}^k U_j = \emptyset$  then set  $Begin_L(q_{i+1}, R)$  and  $End_L(q_{i+1}, R)$  to be  $\emptyset$ . Otherwise set  $Begin_L(q_{i+1}, R)$  to be the highest point on the highest nonempty  $U_j$  ( $0 \leq j \leq k$ ), and set  $End_L(q_{i+1}, R)$  to be the rightmost point on  $\bigcup_{j=0}^k U_j$ . That this sub-step takes  $O(k)$  time can be seen by noting that we already know  $Begin_L(\beta_j, R)$  and  $End_L(\beta_j, R)$ , and hence testing whether  $U_j = \emptyset$  takes constant time (by testing whether  $Begin_L(\beta_j, R) = \emptyset$ ).

(4.2) If  $\bigcup_{j=0}^k U_j = \emptyset$  then go to Step (5). Otherwise let  $U_\alpha$  be the highest nonempty  $U_j$  ( $1 \leq j \leq k$ ).

Repeat the following (i)-(iii):

- (i) Let  $v$  be the rightmost point of  $U_\alpha$ . Let  $U_\gamma$  be the highest nonempty  $U_j$  that is below  $U_\alpha$  (i.e.  $\alpha < \gamma \leq k$ ) and has its rightmost point to the right of  $v$ ; if no such  $U_\gamma$  exists then go to Step (5). Locating  $U_\gamma$  can clearly be done in  $O(\gamma - \alpha)$  time.
- (ii) Start at the leftmost point of  $U_\gamma$  and trace it left-to-right until the first point (say,  $w$ ) to the right of  $v$  is reached: stop the scan of  $U_\gamma$  at  $w$  and add edge  $(v, w)$  to  $E(L, R)$ . We "charge" the cost of tracing the portion of  $U_\gamma$  that is to the left of  $v$  to the points so traced (one unit per point traced).
- (iii) Set  $\alpha := \gamma$  and go to (i).

*Note:* Sub-steps (4.1) and (4.2) can be combined; we chose to keep them separate for ease of exposition.



The cost of sub-step (4.2) is  $O(k)$  plus the cost of the "charges" done in (ii). Let us count the overall cost of the charges done in (ii). A point (say,  $u$ ) that gets charged one unit in (ii) will never get charged again in the future, because when executing Step (4) for a future  $q_{j+1}$  ( $i < j$ ),  $u$  will be "shielded" by  $v$ ; i.e.  $u$  will not belong to the  $MAX(Strip_L(\beta, R))$  of any  $\beta$  in  $MAXREV(\{q_1, \dots, q_j\})$ . Thus the cost of all "charges" done in (ii) is  $O(|L|)$ .

(5) Push  $q_{i+1}$  on *STACK*, then set  $i := i+1$ .

To analyze the time complexity of the above procedure, simply observe that  $q_i \in R$  gets pushed on *STACK* exactly once (once such a point  $q_i$  is removed from *STACK*, it cannot belong to the *MAXREV* of the subset of points of  $Q_R$  already scanned, since at least one point of this subset is above it and to its left). Thus the total time taken by the above procedure is  $O(|L| + |R|)$ .  $\square$

**Corollary 1.**  $G(S)$  has  $O(n \log n)$  edges and can be built in  $O(n \log n)$  time, where  $n = |S|$ .

**Proof.** Choose  $|L| = |R| = n/2$ , and let  $f(n)$  denote the maximum number of edges that  $G(S)$  can have. The edge set of  $G(S)$  consists of the (not necessarily disjoint) union of  $E(L, R)$ ,  $CROSS(L, R)$ , and the edge sets of  $G(L)$  and  $G(R)$ . The number of edges in each of  $G(L)$  and  $G(R)$  is at most  $f(n/2)$ . By observations 1 and 2,  $E(L, R)$  and  $CROSS(L, R)$  have at most  $n/2 - 1$  and  $n/2$  edges, respectively. Therefore  $f(n) \leq 2f(n/2) + n - 1$ , and hence  $f(n) = O(n \log n)$ . The  $O(n \log n)$  time bound for constructing  $G(S)$  is by a straightforward divide and conquer, with Lemma 1 giving the needed linear time conquer step.  $\square$

**Observation 3.** There exists an  $S$  such that  $G(S)$  has  $\Omega(n \log n)$  edges.

**Proof.** Let  $g(n)$  denote the number of edges that  $G(S)$  has by our construction. Construct three identical sets of  $n/3$  points each (call them  $S_1, S_2, S_3$ ), each of which individually gives rise to a  $G(S_i)$  that has  $g(n/3)$  edges. Now, stack  $S_1, S_2, S_3$  on top of one another so that the lowest point in  $S_1$  is higher than the highest point in  $S_2$ , the lowest point in  $S_2$  is higher than the highest point in  $S_3$ , and each point of  $S_1$  has same x-coordinate as the corresponding point of  $S_2$  or  $S_3$ . Now, disturb the above situation as follows: shift every point of  $S_1$  to the right by an extremely small amount  $\epsilon$ , and simultaneously shift every point of  $S_2$  to the left by the same amount  $\epsilon$  (the points

in  $S_3$  don't move). Let  $S$  be the set of points consisting of the union of the new (shifted)  $S_1$ , the new  $S_2$ , and  $S_3$ . The slight shifting of  $S_1$  to the right and  $S_2$  to the left means that for each point  $x_1$  of  $S_1$ , the corresponding point of  $S_2$  (call it  $x_2$ ) is to its left by a  $2\epsilon$  amount, and the corresponding point of  $S_3$  (call it  $x_3$ ) is to its left by an  $\epsilon$  amount. Thus in  $G(S)$ , each  $x_1$  causes the edge  $(x_2, x_3)$  to be present. Thus  $G(S)$  has at least  $3g(n/3) + n/3$  edges, and hence  $g(n) \geq 3g(n/3) + n/3$ , resulting in  $g(n) = \Omega(n \log n)$ .  $\square$

Let the *label* of a point  $p \in S$  with respect to set  $S$  (henceforth denoted  $label_S(p)$ ) be as in the definition of problem P2.

Let  $S$  be partitioned into four subsets  $A_1, A_2, A_3, A_4$ , where  $A_i$  is to the left of  $A_{i+1}$ . For every  $p \in A_{i+1}$ , let  $Left_{A_i}(p, A_{i+1})$  be the smallest  $label_S(q)$  over all  $q$  that are on the portion of  $MD_S(p)$  that lies in  $A_i$ ; that is,

$$Left_{A_i}(p, A_{i+1}) = \text{Min} \{label_S(q) : q \in MAX(Strip_{A_i}(p, A_{i+1}))\} \quad \text{if } Strip_{A_i}(p, A_{i+1}) \neq \emptyset,$$

$$Left_{A_i}(p, A_{i+1}) = \infty \quad \text{otherwise.}$$

**Observation 4.** Let  $p \in A_2$ . If  $DOM_S(p) \neq \emptyset$ , then

$$label_S(p) = w(p) + \text{Min} \{Left_{A_1}(p, A_2), \text{Min} \{label_S(q) : q \in MD_{A_2}(p)\}\}.$$

**Proof.** An immediate consequence of the definitions and the fact that  $MD_S(p)$  is the concatenation of  $MAX(Strip_{A_1}(p, A_2))$  with  $MD_{A_2}(p)$ .  $\square$

**Observation 5.** For every  $p \in A_3$ , we have

$$Left_{A_1 \cup A_2}(p, A_3) = \text{Min} \{Left_{A_1}(p, A_2 \cup A_3), Left_{A_2}(p, A_3)\}.$$

**Proof.** An immediate consequence of the fact that  $MAX(Strip_{A_1 \cup A_2}(p, A_3))$  is the concatenation of  $MAX(Strip_{A_1}(p, A_2 \cup A_3))$  with  $MAX(Strip_{A_2}(p, A_3))$ .  $\square$

### 3. A Special Class of Tree Computations

Chazelle [C] has given a general technique which, given any  $n$  paths on a free tree that has

a real label associated with each node, computes the smallest label on each of these  $n$  paths in  $O(n \log n)$  time. In our algorithm for solving problem P2 (given in the next section), we will need a similar computation on a rooted tree in which the  $n$  paths have a *nested property* (defined below). In Lemma 2, we establish that this can be done in  $O(n)$  time.

**Definition 1.** Let  $C=(P_1, \dots, P_l)$  be a sequence of descendent-to-ancestor paths in a rooted tree  $T$ ; path  $P_i$  begins at  $u_i$  and ends at  $w_i$ , where  $w_i$  is an ancestor of  $u_i$ . We say that  $C$  has the *nested property* iff

- (i)  $i < j$  and  $P_i \cap P_j \neq \emptyset$  imply that  $w_j$  is ancestor of  $w_i$ , and
- (ii)  $i < j < k$  and  $P_i \cap P_j \cap P_k \neq \emptyset$  imply that  $P_i \cap P_k \subseteq P_j \cap P_k$ .

For example, in the tree shown in Figure 3, if  $P_1=a, b, c$ ,  $P_2=u, v, b, c, d$ ,  $P_3=a, b, c, d, e$ , and  $P_4=w, v, b, c, d, e, f$ , then  $(P_1, P_2, P_4)$  has the nested property but  $(P_2, P_3, P_4)$  does not.

**Lemma 2.** Let  $T$  be an  $n$ -node rooted tree represented by *parent* pointers. In addition to *parent*( $v$ ), each node  $v$  also has a real label  $l(v)$  associated with it. Let  $C=(P_1, \dots, P_n)$  be a sequence of descendent-to-ancestor paths in  $T$ . Let  $f(i)$  be the smallest  $l(v)$  over all  $v$  on path  $P_i$ . If  $C$  has the nested property, then  $f(1)f(2), \dots, f(n)$  can be computed in  $O(n)$  time.

**Proof.** We use the path compression technique previously used to solve the UNION-FIND problem [AHU]; the nested property will be crucial in proving that the algorithm actually runs in linear time. Assign to each node  $p$  of  $T$  a temporary label  $Temp(p)$ , initially set to  $l(p)$ ; the significance of these  $Temp$  labels is that as we do path compression on  $T$ , the  $f(i)$  of every  $P_i$  yet to be traced equals the smallest  $Temp$  label on it (this is certainly true initially, and will be maintained as we do path compressions). In what follows, we use  $T_0$  to denote the initial (i.e. unmodified) tree  $T$ , and we view a path  $P_i$  as being defined by its two endpoints  $u_i$  and  $w_i$  rather than by a sequence of nodes in  $T_0$  (path compression on  $T$  may shorten a path in  $T$  but does not change its endpoints).

We process the  $n$  paths in the order  $P_1, \dots, P_n$ . To process  $P_i$ , we first trace it on  $T$  and compute the smallest  $Temp(q)$  over all  $q$  on it, which is  $f(i)$ . Then we modify  $T$  by doing path

compression along the path just traced, as follows. First, by tracing  $P_i$  once in the backward direction (from  $w_i$  to  $u_i$ ), we compute for all  $p \in P_i$ , the quantity  $g(p) = \text{Min} \{ \text{Temp}(q) : q \text{ is on the path from } w_i \text{ to } p \}$ . Once this is done, we modify  $T$  by making every  $p \in P_i - \{w_i\}$  a child of  $w_i$ , and changing its temporary label by doing  $\text{Temp}(p) := \text{Min} \{ \text{Temp}(p), g(p) \}$ . Figure 4 illustrates the effect of this on  $T$  if  $P_i = a, b, c, d$  (in that figure, the numbers between parentheses are  $\text{Temp}$  values).

A  $P_j$  yet to be processed (i.e. one with  $j > i$ ) may have been "shortened" by the path compression made along  $P_i$ ; however, because of property (i) (of Definition 1) and because of the way the  $\text{Temp}$  labels are updated,  $f(j)$  is still the smallest  $\text{Temp}$  on the  $u_j$ -to- $w_j$  path in the modified tree  $T$ . This modification of  $T$  maintains the nested property for the sequence of paths yet to be processed, i.e. for the sequence  $(P_{i+1}, \dots, P_n)$  where every  $P_j$  ( $j > i$ ) is the  $u_j$ -to- $w_j$  path in the modified tree  $T$ ; to see this, observe that every such  $P_j$  ( $j > i$ ) ends at a  $w_j$  whose *parent* pointer is the same as the one in  $T_0$  (because of property (i) and the order in which we are processing the  $P_i$ 's). We now must show that the sum of the lengths of all the  $P_i$ 's traced in this manner is  $O(n)$ . We say that an edge  $e$  of  $T_0$  is *first traced* by  $P_j$  iff  $e$  belongs to  $P_j$  but not to any other  $P_k$  with  $k < j$ . When we trace  $P_i$  in the path-compressed tree  $T$  that resulted from processing paths  $P_1, \dots, P_{i-1}$ , we partition the cost of tracing  $P_i$  into two components: The *strict cost* is that of traversing the edges first traced by  $P_i$ , and the *extra cost* is that of tracing the other edges (the latter may include edges first traced by  $P_j$ 's with  $j < i$  as well as edges previously added by the path compression process). The sum of the strict costs of all the  $P_i$ 's is trivially  $O(n)$ . We now prove that the total extra cost is also  $O(n)$ . Let  $C_i$  denote the set of paths that were processed before  $P_i$  and have a nonempty intersection with  $P_i$  in  $T_0$ , i.e.  $C_i = \{ P_j : j < i \text{ and } P_j \cap P_i \neq \emptyset \text{ in } T_0 \}$ . Let  $P_a$  and  $P_b$  be paths in  $C_i$ ; we say that  $P_a$  *beats*  $P_b$  iff  $a > b$  and  $P_a \cap P_b \neq \emptyset$  in  $T_0$ . Note that if  $P_a$  beats  $P_b$  in  $C_i$ , then the nested property implies that  $P_b \cap P_i \subseteq P_a \cap P_i$  in  $T_0$ . For every  $P_j \in C_i$ , let  $C_{ij}$  denote the subset of  $C_i$  each of whose elements has a nonempty intersection with  $P_j$  in  $T_0$  (see Figure 5). The nested property implies that,

for every  $P_k \in C_{ij}$ ,  $P_{\max(j,k)}$  beats  $P_{\min(j,k)}$ . Path  $P_j$  is said to be a *chief* in  $C_i$  iff it beats every  $P_k \in C_{ij}$ , i.e. iff  $j > \max\{k : P_k \in C_{ij}\}$ . In Figure 5, the chiefs in  $C_i$  are  $P_a$  and  $P_c$ . Let  $D_i$  be the subset of  $C_i$  that contains only the chiefs. The extra cost of tracing  $P_i$  in  $T$  is equal to  $|D_i|$ , because the path compression that was done after processing each chief in  $C_i$  has reduced the intersection of that chief with  $P_i$  to exactly one edge; we "charge" a unit of this extra cost to each chief. A chief in  $C_i$  (say,  $P_a$ ) will be prevented by  $P_i$  from ever being chief in a subsequent  $C_j$  ( $j > i$ ); to see this, note that if such a  $P_j$  ( $j > i$ ) intersects  $P_a$  in  $T_0$  then it must also intersect  $P_i$  in  $T_0$  (because  $a < i < j$ ), and therefore  $P_i$  will belong to  $C_{ja}$  and will beat  $P_a$  in  $C_j$ . Hence the overall extra cost is at most  $n$ .  $\square$

#### 4. Computing the $label_S(p)$ 's

In this section we give an  $O(n \log n)$  time,  $O(n)$  space algorithm for solving problem P2.

Let  $S = \{p_1, \dots, p_n\}$  be the set of input points whose  $label_S(p)$ 's we wish to compute. To simplify the notation, we assume that the  $p_i$ 's are given already sorted by increasing  $x$  coordinates, i.e.  $X(p_1) < X(p_2) < \dots < X(p_n)$ . The algorithm that follows omits the computation of  $predecessor_S(p)$  (including it would have unnecessarily cluttered the exposition). The interested reader can easily modify the algorithm so that it computes  $predecessor_S(p)$  as well as  $label_S(p)$  for all  $p \in S$ . The algorithm is initially called with  $R = S$  and  $Left_{\emptyset}(p, S) = \infty$  for all  $p \in S$ , and it returns with  $label_S(p)$  computed for all  $p \in S$ .

##### Algorithm MAXDOM(R)

*Input:* A contiguous  $m$ -subset  $R$  of  $S$ , i.e.  $R = \{p_r, \dots, p_{r+m-1}\}$ ; for every  $p \in R$ ,  $Left_L(p, R)$ , where  $L = \{p_1, \dots, p_{r-1}\}$ . In addition, the input includes the list  $Q_R$  containing the points of  $R$  sorted by increasing  $y$  coordinates.

*Output:* The labels  $label_S(p_r), \dots, label_S(p_{r+m-1})$ .

*Overview of Algorithm:* The algorithm partitions  $R$  into subsets  $A$  and  $B$  such that  $|A| = |B| = m/2$  and  $A$  is to the left of  $B$ . Since  $Left_L(p, A)$  is given for all  $p \in A$  (it equals

$Left_L(p, R)$ ), the algorithm can recursively call itself for set  $A$ , obtaining  $label_S(p)$  for every  $p \in A$ . Then, using the labels so computed, the algorithm computes  $Left_{L \cup A}(p, B)$  for every  $p \in B$ , in linear time. After that, the algorithm recursively calls itself for set  $B$  (it can do so because it now knows  $Left_{L \cup A}(p, B)$  for all  $p \in B$ ). The trick is how to compute  $Left_{L \cup A}(p, B)$  for all  $p \in B$  in linear time, knowing  $Left_L(p, R)$  for every  $p \in R$  and  $label_S(p)$  for every  $p \in A$ ; lemmas 1 and 2 are used for achieving this.

**Step 1.** If  $m=1$  then set  $label_S(p_r) := w(p_r) + Left_L(p_r, R)$  if  $Left_L(p_r, R) \neq \infty$ ; set  $label_S(p_r) := w(p_r)$  if  $Left_L(p_r, R) = \infty$ . Then return. If  $m > 1$  then proceed to Step 2.

**Step 2.** Let  $A = \{p_r, \dots, p_{r+m/2-1}\}$ ,  $B = \{p_{r+m/2}, \dots, p_{r+m-1}\}$ . Extract from  $Q_R$  the lists  $Q_A$  and  $Q_B$  containing the points of  $A$  and  $B$ , respectively, sorted by increasing  $y$  components.

Step 2 takes  $O(m)$  time.

**Step 3.** Since we have  $Q_A$  and  $Left_L(p, A)$  for every  $p \in A$  (it equals  $Left_L(p, R)$ ), we can recursively solve the problem for the set  $A$  by doing  $MAXDOM(A)$ . This recursive call returns  $label_S(p_r), \dots, label_S(p_{r+m/2-1})$ .

**Step 4.** This step computes the forest  $F = (A, E(A, B))$  together with  $Begin_A(p, B)$  and  $End_A(p, B)$  for every  $p \in B$ . By Lemma 1, this can be done in  $O(m)$  time.

**Step 5.** Let  $Q_B = (b_1, \dots, b_{m/2})$  where  $Y(b_1) < \dots < Y(b_{m/2})$ , and let  $Path(b_i)$  denote the path from  $Begin_A(b_i, B)$  to  $End_A(b_i, B)$  in  $F$ . Use the forest  $F = (A, E(A, B))$  created by the previous step to compute, for every  $p \in B$  such that  $Begin_A(p, B) \neq \emptyset$ , the quantity  $Left_A(p, B) = \text{Min} \{label_S(q) : q \in Path(b_i)\}$ . Lemma 3 (given at the end of this section) shows that the sequence of paths  $Path(b_1), \dots, Path(b_{m/2})$  has the nested property. This and Lemma 2 imply that Step 5 can be done in  $O(m)$  time.

**Step 6.** For every  $p \in B$ , set  $Left_{L \cup A}(p, B) := \text{Min} \{Left_L(p, R), Left_A(p, B)\}$ .

This step takes  $O(m)$  time, and its correctness follows from Observation 5 (in Section 2).

**Step 7.** Recursively solve the problem for set  $B$  by doing  $MAXDOM(B)$ . This returns  $label_S(p_{r+m/2}), \dots, label_S(p_{r+m-1})$ .

**(End of Algorithm)**

**Theorem 1.**  $MAXDOM(S)$  returns  $label_S(p)$  for every  $p \in S$  (and thus solves problem P2) in  $O(n \log n)$  time and  $O(n)$  space.

**Proof.** The running time  $T(m)$  of procedure  $MAXDOM$  satisfies the recurrence  $T(m) \leq 2T(m/2) + O(m)$  and hence  $T(m) = O(m \log m)$ . The space  $S(m)$  satisfies the recurrence  $S(m) \leq S(m/2) + O(m)$ , and thus  $S(m) = O(m)$ . Correctness is easily established by induction on  $|R|$ , using observations 4 and 5.  $\square$

**Lemma 3.** The sequence  $Path(b_1), \dots, Path(b_{m/2})$  of descendent-to-ancestor paths in  $F$  has the nested property.

**Proof.** We first prove property (i) of Definition 1. Let  $i < j$  and assume that  $Path(b_i) \cap Path(b_j) \neq \emptyset$ . Since  $j > i$ ,  $b_j$  is above  $b_i$ . If  $b_j$  were to the right of  $b_i$  then the intersection of  $Path(b_i)$  with  $Path(b_j)$  would be empty, hence  $b_j$  must be to the left of  $b_i$ . Therefore  $leader_B(b_j)$  is not above  $leader_B(b_i)$ . This, and the fact that  $Path(b_i) \cap Path(b_j) \neq \emptyset$ , imply that  $Y(End_A(b_j, B)) \leq Y(End_A(b_i, B))$ . Hence  $End_A(b_j, B)$  is an ancestor of  $End_A(b_i, B)$ . We now prove that property (ii) of Definition 1 also holds. Let  $i < j < k$  and assume that  $Path(b_i) \cap Path(b_j) \cap Path(b_k) \neq \emptyset$ . Property (i) implies that  $End_A(b_k, B)$  is ancestor of  $End_A(b_j, B)$ , which is itself ancestor of  $End_A(b_i, B)$ . Because  $b_i$  is below  $b_j$ , which is below  $b_k$ , we also have

$$Y(Begin_A(b_i, B)) \leq Y(Begin_A(b_j, B)) \leq Y(Begin_A(b_k, B)).$$

This implies that the first (i.e. geometrically highest) point on  $Path(b_i) \cap Path(b_k)$  is an ancestor of the first point on  $Path(b_j) \cap Path(b_k)$ .  $\square$

## 5. Computing the $MD_S(p)$ 's

In this section we briefly sketch how the algorithm of the previous section can be modified to solve problem P1. This problem is considerably easier than P2, and the algorithm (given below) correspondingly simpler.

**Algorithm MD\_LIST**

*Input:* A set  $S$  containing the points  $p_1, \dots, p_n$  where  $X(p_1) < X(p_2) < \dots < X(p_n)$ .

*Output:* The lists  $MD_S(p_1), \dots, MD_S(p_n)$ , together with the list  $Q_S$  containing the points of  $S$  sorted by increasing  $y$  coordinates.

**Step 1.** If  $n=1$  then output  $MD_S(p_1)=\emptyset$  and return. If  $n > 1$  then proceed to Step 2.

**Step 2.** Recursively solve the problem for the set  $A=\{p_1, \dots, p_{n/2}\}$ . This recursive call returns  $MD_S(p_1), \dots, MD_S(p_{n/2})$ , together with the list  $Q_A$  containing the points of  $A$  sorted by increasing  $y$  coordinates.

**Step 3.** Recursively solve the problem for the set  $B=\{p_{n/2+1}, \dots, p_n\}$ . This recursive call returns  $MD_B(p_{n/2+1}), \dots, MD_B(p_n)$ , together with the list  $Q_B$  containing the points of  $B$  sorted by increasing  $y$  coordinates.

*Note.* For every  $p \in B$ , the list  $MD_S(p)$  is the concatenation of  $MAX(Strip_A(p, B))$  with the already computed list  $MD_B(p)$ .  $MAX(Strip_A(p, B))$  is the path from  $Begin_A(p, B)$  to  $End_A(p, B)$  in the forest  $F=(A, E(A, B))$ .

**Step 4.** Construct the forest  $F$ , together with  $Begin_A(p, B)$  and  $End_A(p, B)$  for every  $p \in B$ . This is done in  $O(n)$  time (by Lemma 1).

**Step 5.** Use the forest  $F$  created by the previous step to compute, for every  $p \in B$ , the list  $MAX(Strip_A(p, B))$ . This list is obtained by simply tracing the path in  $F$  from  $Begin_A(p, B)$  to  $End_A(p, B)$  (no path compression is needed since we are interested in the paths themselves rather than in some function of them).

**Step 6.** For every  $p \in B$ , compute  $MD_S(p)$  by concatenating  $MAX(Strip_A(p, B))$  with  $MD_B(p)$ . This takes constant time per concatenation, for a total of  $O(n)$  time.



**Step 6.** Merge  $Q_A$  and  $Q_B$  into  $Q_S$  and return. This takes  $O(n)$  time.

**(End of Algorithm)**

Correctness of the above algorithm is easily established by induction on  $n$ . We analyze its time complexity by charging some of the time to the output, and using  $T(n)$  to denote the time not charged to the output. Thus the total time will be  $O(T(n)+t)$  where  $t = \sum_{p \in S} |MD_S(p)|$ . The cost of Step 5 is completely charged to the output, since every  $MAX(Strip_A(p, B))$  is part of  $MD_S(p)$ . Since the cost charged to  $T(n)$  includes  $2T(n/2)$  plus an additional  $O(n)$  time, we have  $T(n) = O(n \log n)$ . Thus we have established the following.

**Theorem 2.** Algorithm  $MD\_LIST$  correctly solves problem P1, and runs in time  $O(n \log n + t)$ .

## 6. Applications

In this section we discuss some problems for which improved algorithms follow from our solution to the maxdominance problems P1 and P2.

### 5.1. Permutation graphs and subsequence problems

For any undirected graph  $G=(V, E)$ , a subset  $H$  of the vertex set  $V$  is called a *dominating set* iff for every  $u \in V$  there exists  $v \in H$  such that  $u$  is adjacent to  $v$ . Set  $H$  is *independent* iff no two vertices in  $H$  are adjacent. The problem of finding a minimum independent dominating set (MIDS for short) is NP-hard for general graphs, however for the class of *permutation graphs* an  $O(n^3)$  time solution was given in [FK], later improved to  $O(n(\log n)^2)$  in [AMU]. We now briefly point out how our solution to problem P2 implies an  $O(n \log n)$  time solution for the MIDS problem.

In [AMU] the MIDS problem is reduced to that of computing a particular subsequence of a sequence of length  $n$ . Given a sequence  $\alpha = a_1 a_2 \cdots a_n$  of numbers, a *subsequence* of  $\alpha$  is a sequence  $\beta = a_{i_1} a_{i_2} \cdots a_{i_k}$  such that  $i_1 < i_2 < \dots < i_k$ . If, in addition,  $a_{i_1} < a_{i_2} < \dots < a_{i_k}$ , then we say that  $\beta$  is an *increasing subsequence* of  $\alpha$ . An increasing subsequence of  $\alpha$  is *maximal* iff it is not a

proper increasing subsequence of any increasing subsequence of  $\alpha$ . A *maximum* increasing subsequence is one of maximum length. Note that a maximum increasing subsequence is also maximal, but that a maximal increasing subsequence may not be maximum. For example, in the sequence 2,1,4,5,3 the increasing subsequence 1,3 is maximal but not maximum (for this example the length of a maximum increasing subsequence is three, e.g. 2,4,5). In [AMU] it was pointed out that MIDS can be reduced to the problem of computing a *shortest maximal increasing subsequence* (from now on called SMIS) of a sequence of  $n$  numbers. We now point out how our solution to problem P2 implies an  $O(n \log n)$  time solution to the SMIS (and hence MIDS) problem. For the sake of generality, we consider the weighted version of the problem, i.e. where every element  $a_i$  has an associated weight  $w_i$ , and the problem is then to compute a minimum-weight maximal increasing subsequence of the input sequence  $\alpha = a_1 \cdots a_n$ . This is done as follows: create a set of points  $S = \{p_1, \cdots, p_n\}$  where  $p_i = (i, a_i)$ , and let the weight  $w(p_i)$  of point  $p_i$  be  $w_i$ . Let the *label* of every point in  $S$  be defined as follows:

$$\text{label}(p) = w(p) \quad \text{if } \text{DOM}_S(p) = \emptyset,$$

$$\text{label}(p) = w(p) + \text{Min} \{ \text{label}(q) : q \in \text{MD}_S(p) \} \quad \text{otherwise.}$$

As in P2, the predecessor of point  $p$  is any of the points which gave  $p$  its label. It is not hard to see that (i) the minimum-weight shortest maximal increasing subsequence of  $\alpha$  has a weight equal to  $\text{Min} \{ \text{label}(p) : p \in \text{MAX}(S) \}$ , and (ii) the corresponding subsequence of  $\alpha$  can be retrieved by beginning at the smallest-labeled point in  $\text{MAX}(S)$  and following the chain of predecessor pointers. These observations imply that our solution to problem P2 implies a solution to SMIS (and hence MIDS) having complexity  $O(n \log n)$  time and  $O(n)$  space.

The known  $O(n \log n)$  time solutions to the well studied problem of computing a maximum increasing subsequence [D,DMS] cannot be modified to solve the SMIS problem, which is considerably more difficult in spite of the apparent similarity.

## 5.2. Empty rectangle problem

Given a rectangle  $R$  and a set  $S$  of  $n$  points in  $R$ , a *valid rectangle* is one which is contained in  $R$ , has its sides parallel to those of  $R$ , and does not contain any of the points in  $S$ . Consider the problem of enumerating all the *restricted rectangles*, where a *restricted rectangle* (RR for short) is a valid rectangle such that each of its four edges either contains a point of  $S$  or coincides with an edge of  $R$ . Let  $s$  denote the number of RR's, i.e. the size of the output. Naamad et al. [NLH] prove that  $s=O(n^2)$  and give an example in which  $s=\Theta(n^2)$ . They also show that when the points are drawn from a uniform distribution, the expected value of  $s$  is  $O(n \log n)$ .

In [AF] it was shown that any  $O(T(n)+t)$  time algorithm for problem P1 would immediately imply an  $O(T(n)+s)$  time algorithm for enumerating all the RR's (recall that  $t$  is the size of the output to P1). The solution that was given in [AF] had  $T(n)=n(\log n)^2$ . Since our solution to P1 has  $T(n)=n \log n$ , it automatically implies an  $O(n \log n + s)$  time solution to the problem of enumerating all RR's. This is an improvement over the  $O(n(\log n)^2 + s)$  time algorithm given in [AF] and over the  $O(\min(n^2, s \log n))$  time algorithm given in [NLH].

Since the expected value of  $s$  is  $O(n \log n)$ , our result implies an improvement by a factor of  $\log n$  in the best known average case time complexity for the related problem of computing the largest (i.e. maximum area) RR. Similar bounds (using a different method) were recently independently established in [BE,O,PR]. A worst-case time bound of  $O(n(\log n)^3)$  for finding the largest RR was given in [CDL], recently improved to  $O(n(\log n)^2)$  in [AS].

### 5.3. Independent Sets in Overlap or Circle Graphs

Given  $n$  intervals  $I_1, \dots, I_n$  on the line, their corresponding *overlap graph* is the undirected graph having the  $I_i$ 's as vertices, and such that there is an edge between intervals  $I_i$  and  $I_j$  iff these two intervals overlap but neither one contains the other. The problem of computing a maximum-weight independent set for such graphs (which are the same as circle graphs) was considered in [AH] and an algorithm of time complexity  $O(nd)$  was given, where  $d \leq n$  is a quantity whose expected value is proportional to  $n$ . Thus the average case time complexity of the algorithm given in [AH] is  $O(n^2)$ . Our solution to problem P1 makes possible an  $O(n \log n + t)$  time

implementation of the same algorithm that was given in [AH], where  $t$  can still be quadratic in the worst case but has expected value  $O(n \log n)$  (the algorithm is essentially the same as that of [AH] and we therefore refer the interested reader to that paper).

## 7. Conclusion

We gave asymptotically optimal algorithms for two maxdominance problems. These in turn implied improvements in the time complexities of a number of graph-theoretic and geometric problems. The techniques we used are of independent interest, and we have reason to believe they will be useful for solving other problems as well.

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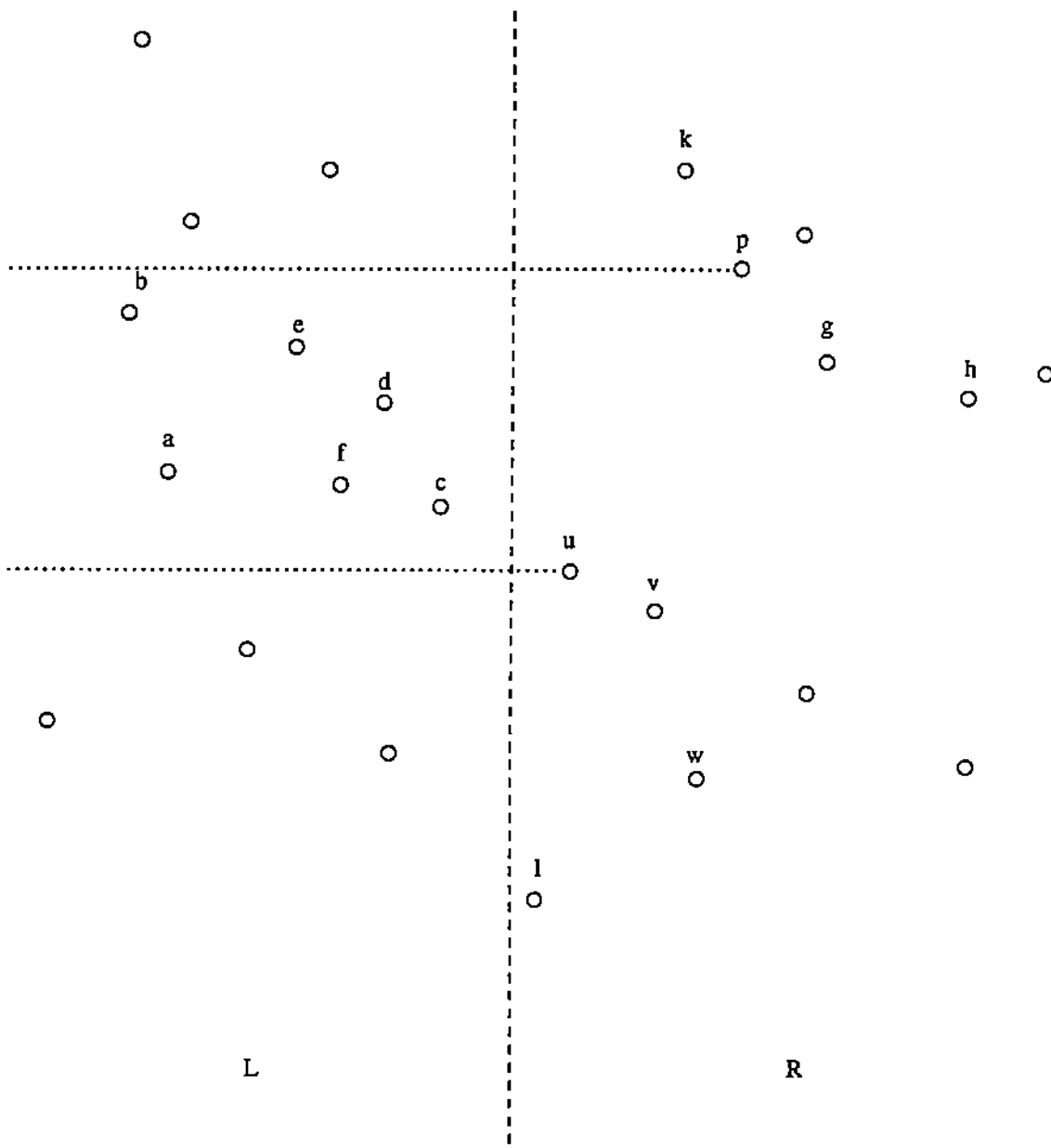


Figure 1. Illustrating the basic definitions

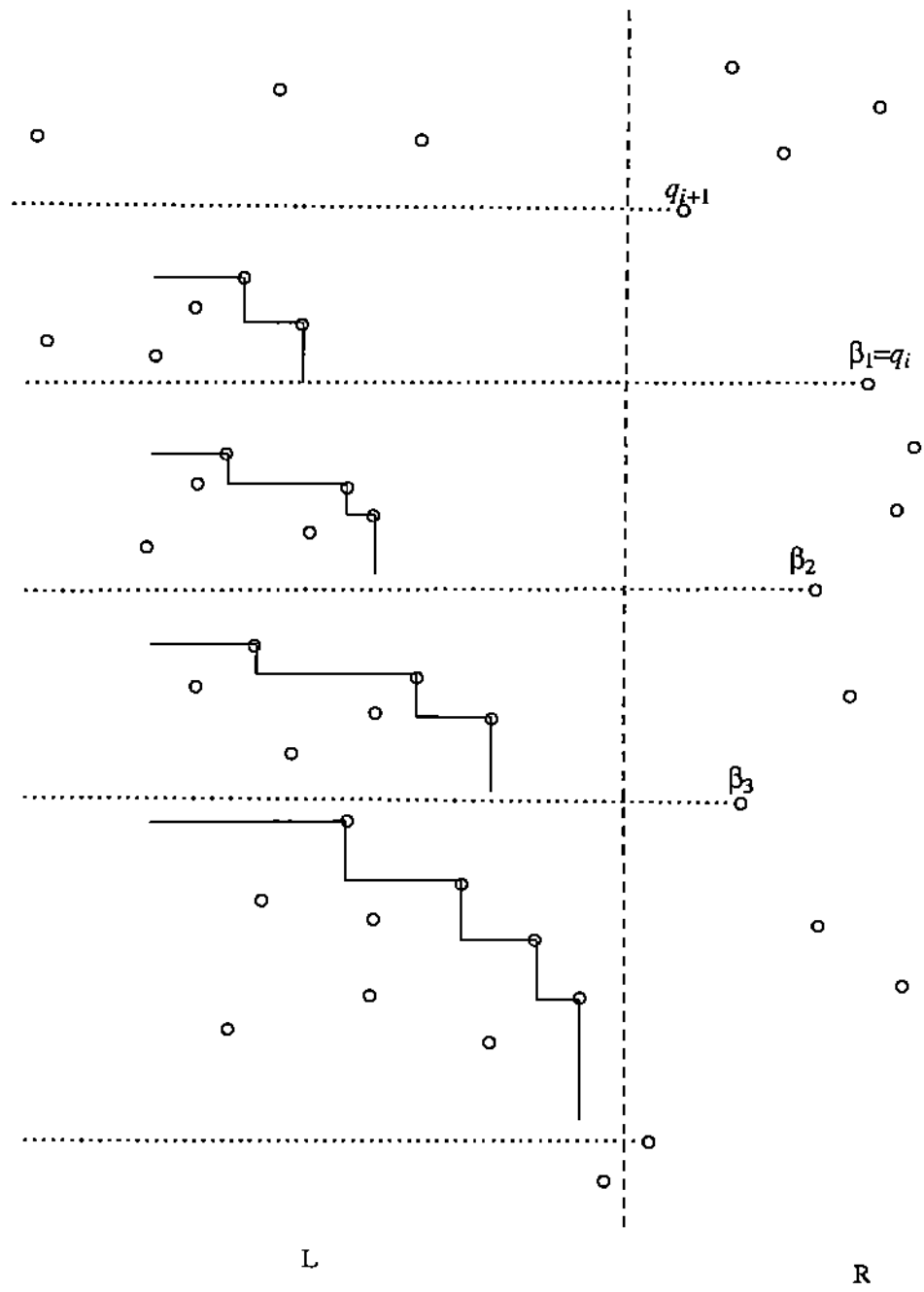


Figure 2. Illustrating Step 4 of the proof of Lemma 1.

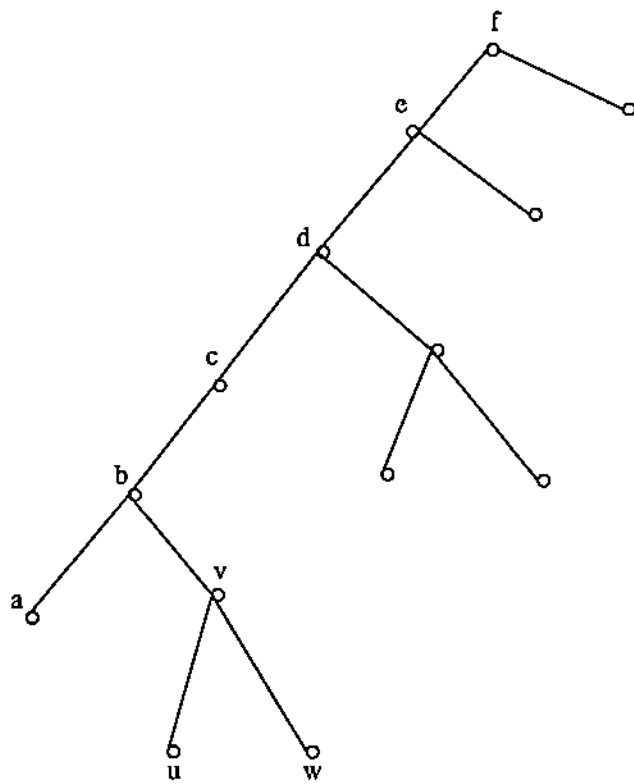


Figure 3. Illustrating Definition 1.



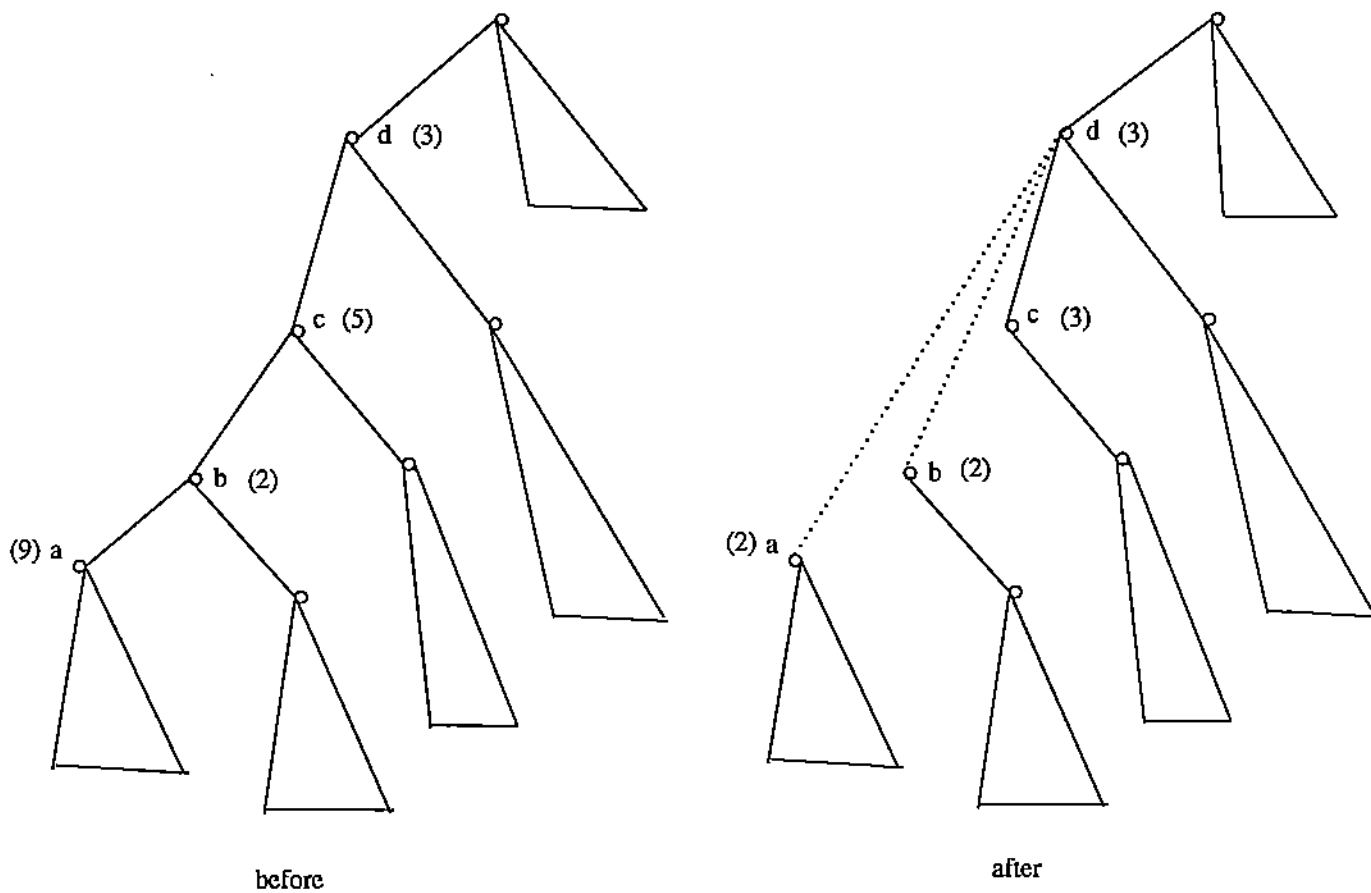


Figure 4. Illustrating how the *Temp* labels are updated.

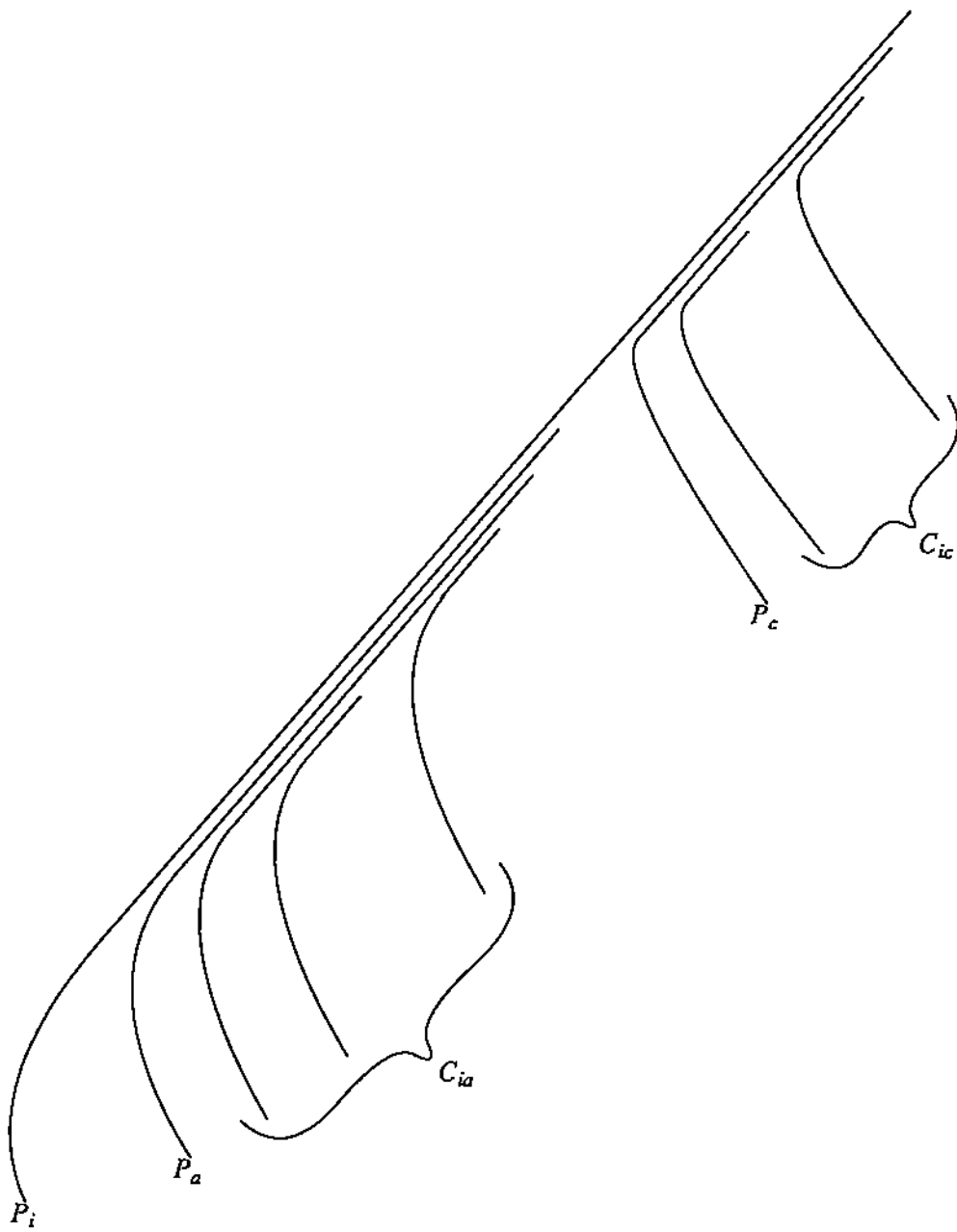


Figure 5. Here  $C_i$  contains seven paths. All paths shown are in  $T_0$ .