Algorithm for the Length-Constrained Maximum-Density Path Problem in a Tree with Uniform Edge Lengths**

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1. Introduction

Let $T$ be an edge-weighted tree with $n$ vertices. Each edge $e$ in $T$ is associated with a real number, called its weight, $w_e$. Let $\pi_{u,v}$ denote the path connecting two distinct vertices $u, v$ in $T$. The weight of $\pi_{u,v}$ is the sum of the weights of the edges in it, $w(\pi_{u,v}) = \sum_{e \in \pi_{u,v}} w_e$, its length $l(\pi_{u,v})$ is the number of edges in it, and its density is $d(\pi_{u,v}) = w(\pi_{u,v})/l(\pi_{u,v})$.

The length-constrained maximum-density path problem is: Given an edge-weighted tree $T$ and a positive integer $L$, find a maximum-density path of length at least $L$ in $T$, namely, a path $\pi$ such that $d(\pi) = \max\{d(\pi_{u,v}) | u,v \text{ are vertices in } T \}$ and $l(\pi_{u,v}) \geq L$.

For this problem, Lin, Kuo, and Chao [11] presented an $O(nL)$ time dynamic programming algorithm, and Lau, Ngo, and Nguyen [10] gave an $O(n \log n)$ time, divide-and-conquer algorithm based on centroid decomposition. In this paper, we develop an $O(n)$ time algorithm.

In Sect. 2, we review some previous results that will be used in designing our algorithm. In Sect. 3, our algorithm for the case where $T$ is a binary tree is first presented, and then we show how to tackle the case of general trees.

A similar approach was previously used to solve another problem by the author [9].

** Manuscript received January 8, 2014.
Manuscript revised March 25, 2014.
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**This research was supported by Basic Science Research Program through the National Research Foundation of Korea (NRF) funded by the Ministry of Education, Science and Technology (No. 2012008222) and by the Chung-Ang University research grant in 2013.
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DOI: 10.1587/transinf.2014EDP7009

2. Preliminaries

The following lemma was first presented by Huang [7] and was previously used in [11], which enables us to consider only the paths of length at least $L$ and at most $2L - 1$ in finding a length-constrained maximum-density path.

Lemma 1: If $\pi$ is a length-constrained maximum-density path, then $l(\pi) \leq 2L - 1$.

Chung and Lu [2] presented a linear time algorithm for the following problem: Given an array $A$ of $n$ real numbers and two positive integers $L \leq U$, find a maximum-density segment $A[j, j']$ with $L \leq j' - j + 1 \leq U$. This corresponds to an array version of the length-constrained maximum-density path problem with both lower and upper length bounds**. They also gave a linear time algorithm for a “location-constrained” variant of the problem in which two integer intervals $[x, y]$ and $[x', y']$ with $y < x'$ are additionally given and it is required that $j \in [x, y]$ and $j' \in [x', y']$.

Let $T$ be a binary tree with $n$ vertices. A component of $T$ is a connected subtree of $T$. Two components are adjacent if there is an edge of $T$ such that one of its endpoints is in one component and the other is in the other. The external degree of a component is the number of its adjacent components.

Frederickson [5] developed a method that partitions a binary tree into a number of components by removing some of its edges and builds a hierarchical sequence of partitions based on incremental merging of the components. Eppstein [3], [4], and Italiano and Ramaswami [8] used the method in developing algorithms for dynamically maintaining trees and graphs. We borrow the definitions from [8] with slight modification.

For an integer $2 \leq z \leq n$, a restricted partition of order $z$ of $T$ is a partition of $T$ into components such that

1. Each component has external degree at most three.
2. Each component of external degree three consists of a single vertex.
3. Each component of external degree one or two consists of at most $z$ vertices.
4. No two adjacent components can be combined and still satisfy 1–3.

Figure 1 shows a restricted partition of order two of a

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The following gives a hierarchical method of reconstructing $T$ from the components by restoring the removed edges, which has certain nice properties (detailed later). A restricted multilevel partition of order $z$ is a hierarchical sequence of partitions of $T$ such that:

1. The components at level 1 (called basic components) are those of a restricted partition of order $z$ of $T$.
2. Each component at level $i > 1$ is either a copy of a component at level $i - 1$ or is formed by combining two adjacent components at level $i - 1$ with the addition of the edge between them. This is done by finding a restricted partition of order two of the tree obtained after contracting the components at level $i - 1$.
3. There is only one component, $T$, at the topmost level.

Figure 2 depicts a tree and its restricted multilevel partition of order two.

A rooted binary tree, called a topology tree for $T$, can naturally be formed from a restricted multilevel partition of $T$.

1. A node at level $i$ in the topology tree represents a component at level $i$ in the restricted multilevel partition.
2. A node at level $i > 1$ has at most two children that represent the components at level $i - 1$ whose union is the component represented by the node.

Frederickson [5] proved that

1. A restricted partition of order $z$ of $T$ can be found in $O(n)$ time and the number of components in the partition is $O(n/z)$.
2. A topology tree has $O(n/z)$ leaf nodes and is of height $O(\log(n/z))$, and
3. a topology tree can be constructed in $O(n)$ time.

Figure 3 illustrates the topology tree for the tree in Fig. 2.

A solution for the following path sum query problem is employed in the design of our algorithm. Let $T$ be a tree in which the edges are associated with real values. The path sum of the path $\pi_{uv}$ for vertices $u, v$ in $T$, denoted by $s(\pi_{uv})$, is the sum of the values of the edges in $\pi_{uv}$. In our application, $s(\pi_{uv})$ is $l(\pi_{uv})$ or $w(\pi_{uv})$. The path sum query problem is: Preprocess $T$ so that, after preprocessing, for any query pair of vertices $u, v$ in $T$, the path sum $s(\pi_{uv})$ can be answered quickly. A solution for the problem consists of two parts: preprocessing and query-answering.

The following solution is well-known, whose preprocessing works as follows:

1. Pick up a vertex $t$ in $T$ and convert $T$ into a rooted tree with root $t$.
2. Compute $s(\pi_{tv})$ for each vertex $v$ in $T$.
3. Preprocess $T$ so that, given any query pair of vertices $u, v$ in $T$, their lowest common ancestor can be found quickly.

After the preprocessing, given any pair of vertices $u$ and $v$, the query-answering finds their lowest common ancestor $t'$ and computes $s(\pi_{uv}) = s(\pi_{ut'}) + s(\pi_{t'v}) - 2 \cdot s(\pi_{tt'})$. Since solutions with $O(n)$ preprocessing time and $O(1)$ query answering time for the lowest common ancestor problem are given by Harel and Tarjan [6], and Schieber and Vishkin [12], the path sum query problem can be solved within the same complexity.

3. Algorithm

Let $T$ be an edge-weighted binary tree with $n$ vertices for which we want to solve the length-constrained maximum-density path problem. Let $L$ be a positive integer that corresponds to lower length bound.

Briefly, our algorithm works as follows: We first divide $T$ into components by removing some of the edges of $T$, and separately solve the problem for each of the components. These components and their solutions are at level 1. To reach the topmost level where only one component $T$ itself and its solution exist, we explain how to go from level $i - 1$ to level $i$ for $i > 1$. We select some distinct pairs of adjacent components out of those at level $i - 1$. For each pair selected, we merge the two components by restoring
the edge between them and combine their solutions at level $i - 1$ to get one at level $i$. The unselected components and their solutions at level $i - 1$ are copied and stored at level $i$. We repeat this procedure until the topmost level reaches. Described later are details about how large the components at level 1 are, which pairs are selected from the components at level $i - 1$, and how the two solutions at level $i - 1$ are combined to get one at level $i$.

We first perform preprocessing of the path sum query on $T$ with respect to both length and weight of edges so that, after preprocessing, $l(\pi_{u,v})$ and $w(\pi_{u,v})$ for any two vertices $u, v$ in $T$ can be found in $O(1)$ time. This preprocessing takes $O(n)$ time as mentioned in Sect. 2.

We next build in $O(n)$ time a topology tree $T^*$ that corresponds to a restricted multilevel partition of order $z = L$ of $T$. As explained in Sect. 2, $T^*$ is a rooted binary tree of height $O(\log(n/L))$ with $O(n/L)$ leaf nodes, and each node of $T^*$ represents a component of $T$. Let $T_p$ for a node $p$ in $T^*$ denote the component of $T$ that $p$ represents. If $p$ is a leaf, then $T_p$ is a basic component and has at most $L$ vertices. If $p$ has a single child $q$ only, then $T_p = T_q$. If $p$ has two children $q$ and $r$, then $T_p = T_q \cup T_r$, where $(b, c, d)$ is the path between $T_q$ and $T_r$, and $b$ and $c$ are vertices in $T_q$ and $T_r$, respectively. See Fig. 4. In other words, two components $T_q$ and $T_r$ are combined to produce $T_p$ by restoring the edge. We call $b$ (respectively, $c$) the connector of $T_q$ (respectively, $T_r$). If $p$ has a single child $q$, then we define the connector of $T_q$ to be equal to the connector of $T_p$. So, each $T_p$ has one connector.

$T^*$ requires $O(n)$ memory as proved in [5]. We augment $T^*$ by storing the basic component $T_p$ at each leaf node $p$, and the connector of $T_p$ at each non-root node $p$. This augmented $T^*$ clearly requires $O(n)$ memory.

For a node $p$ in $T^*$ and a vertex $a$ in $T$, define $W^a_p$ to be an array of size $2L - 1$ such that, for $0 \leq i \leq 2L - 2$, $W^a_p[i] = \max\{w(\pi_{a,v}) \mid v \text{ is a vertex in } T_p \& l(\pi_{a,v}) = i\}$. (Refer to Lemma 1.) $W^a_p[i]$ is the maximum of the weights of the paths of length $i$ in $T$ between $a$ and the vertices of $T_p$, and $W^a_p[i] = -\infty$ if there is no such path.

If $a$ belongs to $T_p$, then $W^a_p$ can be computed in $O(\max(|T_p|, L))$ time by the following “direct” method. Set $W^a_p[i] = -\infty$ for all $0 \leq i \leq 2L - 2$. Transform $T_p$ into a rooted tree by naming $a$ its root. Traverse $T_p$ in preorder and for each currently visited vertex $v$, set $W^a_p[l(\pi_{a,v})] = \max\{w(\pi_{a,v}), W^a_p[l(\pi_{a,v})]\}$ if $l(\pi_{a,v}) \leq 2L - 2$. Note that $l(\pi_{a,v})$ and $w(\pi_{a,v})$ can be found in $O(1)$ time using the path sum query.

For notational simplicity, $W^a_p$ is used to denote $W^a_p$ if $a$ is the connector of $T_p$. We compute $W^a_p$ for the nodes of $T^*$. If we use the “direct” method for every node $p$, it takes at least $O(n \log(n/L))$ time as $T^*$ has $O(\log(n/L))$ levels and at each level $O(n)$ time is sufficient to apply the “direct” method to the nodes at the level. We show that this can be done in $O(n)$ time by presenting an algorithm that works in bottom-up fashion.

If $p$ is a leaf, then $T_p$, which is a basic component, contains at most $L$ vertices and thus $W^a_p$ can be computed in $O(L)$ time by the “direct” method. If $p$ has a single child $q$ only, then $W^a_p = W^a_q$.

Consider the case where $p$ has two children $q$ and $r$. Both $W^a_q$ and $W^a_r$ have been computed and are available for reference. As mentioned before, $T_p = T_q \cup T_r \cup \{b, c\}$, where $b$ and $c$ are the connectors of $T_q$ and $T_r$, respectively. Let $a$ be the connector of $T_p$. Then, $W^a_p[i] = \max\{W^a_q[i], W^a_r[i]\}$ for $0 \leq i \leq 2L - 2$.

In order to explain how to compute $W^a_q$ and $W^a_r$, assume, without loss of generality, that $a$ belongs to $T_q$. We first compute $W^a_p$ from $W^a_q$. Note that $W^a_p = W^a_q$ if $c$ is the connector of $T_r$. Since $a$ is not in $T_r$, any path from $a$ to a vertex of $T_r$ contains $\pi_{a,c}$ as its subpath (See Fig. 4). Thus, $W^a_p[i] = -\infty$ for $0 \leq i \leq \min(l(\pi_{a,c}) - 1, 2L - 2)$, and $W^a_p[i] = \max\{w(\pi_{a,c}) + W^a_q[i - l(\pi_{a,c})], 2L - 1\} \leq i \leq 2L - 2$. Since both $l(\pi_{a,c})$ and $w(\pi_{a,c})$ can be obtained in $O(1)$ time, $W^a_p$ can be found in $O(L)$ time.

$W^a_q$ can be computed recursively. Both $T_q$ and $T_r$ contain $a$ in Fig. 4. So, computing $W^a_q$ and computing $W^a_q$ are basically identical except that $T_p$ and $T_q$, respectively, are involved. The method that computes $W^a_q$ can be used to compute $W^a_q$ if $p$ is replaced by $q$ (or, if $T_p$ is replaced by $T_q$). The recursion stops if a leaf node is reached. Summarizing the explanations given so far, our algorithm for computing $W^a_p$ is shown in Fig. 5. To compute $W^a_p$ for the nodes of $T^*$, traverse $T^*$ in preorder and call $\text{computeW}(p)$ for each node $p$.

We analyze the time complexity of computing $W^a_p$ for the nodes of $T^*$. Remember that $T^*$ is a rooted binary tree of height $h = O(\log(n/L))$, and has $O(n/L)$ leaf nodes at level 1 and its root at level $h$. Moreover, Frederickson [5].

Fig. 4 The case of $p$ having two children $q$ and $r$.

Fig. 5 Algorithm for computing $W^a_p$. 

computeW($p$)
$\quad a =$ the connector of $T_p$
$\quad$ return recursion($p, a$);

recursion($p, a$)
if ($p$ is a leaf), then compute $W^a_q$ “directly” and return it;
if ($p$ has a single child $q$), then return $W^a_q$;
if $p$ has two children $q, r$ and assume that $a$ is in $T_q$:
compute $W^a_q$ as explained in the text,
$W^a_q =$ recursion($q, a$);
for ($i = 0$ to $2L - 2$)  
$W^a_p[i] = \max\{W^a_q[i], W^a_r[i]\}$;
return $W^a_p$;
proved that the number of nodes at level $i$ is at most $5/6$ of the number of nodes at level $i - 1$. So, if $m$ is the number of leaf nodes of $T^*$, then the number of nodes at level $i$ is at most $(5/6)^i m$.

For a node $p$ at level $i$, a call to $\text{computeW}(p)$ makes at most $i$ recursive calls to $\text{recursion}()$ until a leaf is reached. So, the total number of calls to $\text{recursion}()$ while computing $W_p$ for the nodes $p$ of $T^*$ is at most $\sum_{i=1}^{\infty} i (5/6)^i m$, which is less than $36m$. Note that $\sum_{i=0}^{\infty} i a^i = a/(1 - a)^2$ for $0 < a < 1$. Since each call to $\text{recursion}()$ takes $O(L)$ time and $m = O(n/L)$, the time of computing $W_p$ for the nodes $p$ of $T^*$ is at most $30m \cdot c_1 L \leq 30 \cdot c_2(n/L) \cdot c_1 L = O(n)$ for constants $c_1, c_2 > 0$.

We now have $W_p$ for every node $p$ of $T^*$. To locate a length-constrained maximum-density path in $T$ we need to consider only the paths of length at least $L$ and at most $2L - 1$. Our idea is to traverse $T^*$ in postorder and to find a maximum-density path $\pi_p$ in $T_p$ of length at least $L$ and at most $2L - 1$ and its density $d_p$ for each node $p$ of $T^*$. Then, $d_{\text{root}}$ for the root of $T^*$ is the density of a length-constrained maximum-density path of $T$. For ease of explanation, our algorithm is designed to compute the density $d_p$ only, and an easy modification of the algorithm locates the path $\pi_p$ itself.

If $p$ is a leaf node in $T^*$, then $T_p$ has at most $L$ vertices and thus contains no path of length at least $L$. So, set $d_p = -\infty$. If $p$ has a single child $q$ in $T^*$, then $T_p = T_q$ and so set $d_p = d_q$.

Suppose that $p$ has two children $q$ and $r$. Then, $T_p = T_q \cup T_r \cup \{\hat{e}\}$, where $\hat{e}$ is the edge connecting $T_q$ and $T_r$. At this point, we have $d_q$ and $d_r$. Since $d_q$ (respectively, $d_r$) is the density of a length-constrained maximum-density path of $T_q$ (respectively, $T_r$), we need $d_p$, the density of a length-constrained maximum-density path such that one of its endpoints is in $T_q$ and the other is in $T_r$. Then, $d_p = \max(d_q, d_r, d')$.

To compute $d'$, we employ the linear time algorithm for the “location-constrained” version of the problem of Chung and Lu [2] reviewed in Sect. 2. From $W_q$ and $W_r$, construct an array $A[-2L + 2 : 2L - 2]$ of size $4L - 3$ as follows: $A[0] = w_q, A[-l] = W_q[i] - W_q[i-1]$ for $1 \leq i \leq 2L - 2, \text{and } A[l] = W_r[i] - W_r[i-1]$ for $1 \leq i \leq 2L - 2$. For $j \in [-2L + 2, -1]$ and $j' \in [1, 2L - 2]$, we have $\sum_{i=j}^{j'} A[i] = W_q[-j] + w_q + W_r[j']$, which is the maximum of the weights of the paths of length $-j + j' + 1$ such that one of their endpoints is in $T_q$ and the other is in $T_r$. After setting $U = 2L - 1, [x, y] = [-2L + 2, -1]$, and $[x', y'] = [1, 2L - 2]$, find a solution on $A$ using the algorithm, which is $d'$. This takes $O(L)$ time.

Since computing $d_p$ for a node $p$ of $T^*$ takes $O(L)$ time and there are $O(n/L)$ nodes in $T^*$, the computation of $d_p$ for the nodes $p$ of $T^*$ takes $O(n)$ time.

Our algorithm can be summarized as follows:

1. Perform the preprocessing of the path sum query on $T$ with respect to both edge length and edge weight.
2. Build a topology tree $T^*$ which corresponds to a restricted multilevel partition of order $z = L$ of $T$.
3. Compute $W_p$ for the nodes $p$ of $T^*$ by traversing $T^*$ in postorder.
4. Compute $d_p$ for the nodes $p$ of $T^*$ by traversing $T^*$ in postorder. Then, $d_{\text{root}}$ is the final solution.

**Theorem 1:** A length-constrained maximum-density path in a binary tree can be computed in $O(n)$ time.

To find a length-constrained maximum-density path in a (general) tree, we transform the tree into a binary tree. Given an edge-weighted tree $T_0$, a binary tree $T$ is obtained as follows [13]: Select a vertex of degree one of $T_0$ and make it the root of the rooted version of $T_0$. For each vertex $v$ of $T_0$ with $k \geq 3$ children, $v_1, \ldots, v_k$, replace $v$ with a path $(v, u_2, \ldots, u_{k-1})$, where $(u_2, \ldots, u_{k-1})$ are new vertices and $(v, u_2), (u_2, u_3), \ldots, (u_{k-2}, u_{k-1})$ are new edges. Each of these new edges has $l_e = w_e = 0$. Replace the edges $(v, v_1) | 2 \leq i \leq k - 1$ with the edges $(u_i, v_i) | 2 \leq i \leq k - 1$ of corresponding weights, and replace the edge $(v, v_1)$ with the edge $(u_{k-1}, v_k)$ of corresponding weight. The edge $(v, v_1)$ remains unchanged. Each of these “old” edges has $l_e = 1$. See Fig. 6. A proof that every path in $T_0$ has a corresponding path in $T$ with the same length and weight, and vice versa can be found in [1]. This transformation runs in $O(n)$ time.

Now, $T$ is an edge-weighted binary tree with binary edge lengths. The length of a path is redefined to be the sum of the lengths of the edges in the path, or the number of edges with $l_e = 1$. Theorem 1 is about binary trees with uniform edge lengths (all $l_e = 1$) while the current $T$ is a binary tree with binary edge lengths. We focus on explaining what must be changed in the description of the algorithm above when the edge lengths are not uniform but binary.

Since the preprocessing of the path sum query can be done with respect to any real values, step 1 is not affected by the fact that the edge lengths are binary. Since a restricted multilevel partition and a topology tree are constructed on the unweighted version of $T$, step 2 can be executed as is.

In Fig. 5, when $p$ is a leaf, $W_p$ can be computed using the “direct” method without any change since each $l(\pi_{\text{root}})$ is from the path sum query. Similarly, the computation of $W_p$ is unchanged because the path sum query enables us to find $l(\pi_{\text{root}})$. So, we can execute step 3 as in the algorithm above.

In step 4, the only part that needs explanation is about $A$. We construct an array $A[-2L + 2 : 2L - 2]$ of size $4L - 3$ from $W_q$ and $W_r$ as we did previously. For $j \in [-2L + 2, -1]$ and $j' \in [1, 2L - 2]$, $\sum_{i=j}^{j'} A[i] = W_q[-j] + w_q + W_r[j']$, which is the maximum of the weights of the paths of length

![Fig. 6](image-url)
− j + j′ + l_e such that one of their endpoints is in \( T_q \) and the other is in \( T_r \). If \( l_e = 0 \), then \( w_e = 0 \) and thus \( \sum_{i=j}^{j} A[i] = W_q[j] - j + W_r[j′] \). So, step 4 correctly works.

As each of steps 1–4 again takes \( O(n) \) time, we have proved the following theorem.

**Theorem 2:** A length-constrained maximum-density path in a tree can be computed in \( O(n) \) time.

References


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