Deterministic approximation for the cover time of trees

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Abstract

We present a deterministic algorithm that given a tree T with n vertices, a starting vertex v and a slackness parameter $\epsilon > 0$, estimates within an additive error of ϵ the *cover and return time*, namely, the expected time it takes a simple random walk that starts at v to visit all vertices of T and return to v. The running time of our algorithm is polynomial in n/ϵ, and hence remains polynomial in n also for $\epsilon = 1/n^{O(1)}$. We also show how the algorithm can be extended to estimate the expected cover (without return) time on trees.

1 Introduction

Let G be a connected graph with vertices $v_1, \ldots v_n$. We consider simple random walks on G. Namely, the walk starts at some vertex of the graph, and at every time step picks at random with uniform probability a neighbor of the current vertex and moves to it. Let $C_1^+(G)$ (the *expected cover and return time*) denote the expected number of steps it takes a random walk that starts at v_1 to visit all vertices of G and return to v_1. An empirical estimate for the value of $C_1^+(G)$ can be obtained by starting a random walk at v_1 and counting the number of steps until it visits all vertices of G and returns to v_1. Averaging multiple such estimates one obtains with high probability an accurate approximation for $C_1^+(G)$. An approximation within a multiplicative error of $1 \pm \epsilon$ with probability $1 - \delta$ can be obtained in time polynomial in n, $1/\epsilon$, and $1/\delta$. This follows from the fact that for every graph $C_1^+(G) < n^3$ (see [3, 4, 7]). The question of whether there is a deterministic algorithm that approximates $C_1^+(G)$ within a multiplicative error of $1 \pm \epsilon$ in time polynomial in n and $1/\epsilon$ is open (see for example Chapter 8 in [2]). Here we provide a positive answer to this question in the special case that the underlying graph is a tree.

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1.1 Related work

A survey of random walks in graphs is provided by Lovasz [12]. A related book in preparation by Aldous and Fill [2] is also available on the web. There are also additional books that contain much information on random walks in graphs, such as the recent book by Levin, Peres and Wilmer [11]. More information and appropriate references for some of the well known claims that we make below can be found in these references.

A random walk on a graph is a special case of a Markov chain, with the vertices of the graph serving as the states of the Markov chain, and the edges providing an implicit representation for the transition probabilities. Some parameters of interest for random walks are the *expected hitting time* (expected number of steps it takes to get from one given vertex to another given vertex), the *expected commute time* (expected number of steps to make a round-trip between two given vertices) and the *expected cover time* (the expected number of steps that it takes to visit all vertices). By convention, throughout this paper, we omit the qualifier *expected* when we deal with expectations of random times, and write *hitting time* for the expected hitting time, etc. When dealing with the actual random variables instead of their expectations, we use the term *random hitting time*, etc.

The main result described in this manuscript refers to the *cover and return time*, which requires walks to return to the starting vertex after covering the graph. The hitting time and the commute time can be computed in polynomial time (by solving a system of linear equations). In particular, let us note here that the commute time between any two adjacent vertices in a tree with n vertices is exactly $2(n-1)$. The cover time and cover and return time can be computed in exponential time (again by solving a system of linear equations, but the number of variables is exponential in the size of the graph). It is not known whether there is a polynomial time algorithm for computing either the cover time or the cover and return time.

As noted earlier, there is a natural randomized algorithm that in polynomial time estimates the cover time (or alternatively, the cover and return time), up to some small error. The question of whether in general the use of randomness helps (in a substantial way) in the design of polynomial time algorithms (or in complexity theoretic terms, is BPP=P?) has a natural counterpart in the context of the cover time, namely, can a deterministic polynomial time algorithm achieve as good an approximation of the cover time as the randomized algorithm? This question has been studied in the past, with moderate success.

Much of previous work dealt with the cover time from the worst possible starting vertex in the graph. In this case, the maximum hitting time serves as a lower bound on the cover time. Moreover, as shown by Matthews [13], the cover time can exceed the maximum hitting time by a factor of at most $\ln n$. Hence the hitting time (which is computable in deterministic polynomial time) provides a $\ln n$ approximation to the cover time. An extension of this approach leads to an algorithm with a better approximation ratio of $O((\log\log n)^2)$ [10]. An approach of upper bounding the cover time based on spanning trees is presented in [4]. In particular, when it is applied to trees it implies that the cover and return time is at most $2(n-1)^2$ (which is attained for a path with n vertices), and for general graphs it gives an upper bound of n^3 (which can be improved to essentially $4n^3/27$ with more careful analysis [7]). For some graphs, this approach based on spanning trees gives a very good approximation of the cover time.

When one seeks to estimate the cover time from a given vertex (rather than from the worst possible vertex), the known bounds deteriorate. The deterministic algorithms known [9, 5] pay an extra $O(\log n)$ factor in the approximation ratio compared to the approximation ratios known from worst possible vertex. For the special case of trees, some upper bounds are presented in [8].

There are some special families of graphs for which the cover time is known exactly (e.g., for paths, cycles and complete graphs), or almost exactly (e.g., for balanced trees [1] and for two and higher dimensional grids [6, 2]).

1.2 Our results

Our main theorem is the following.

Theorem 1.1 *There is a deterministic algorithm that given as input a tree T on n vertices, a starting vertex v and a slackness parameter $\epsilon > 0$, outputs a value $A(T, v, \epsilon)$ that approximates the cover and return time $C_v^+(T)$ within a factor of $1 \pm \epsilon$. Namely,*

$$(1-\epsilon)A(T, v, \epsilon) \leq C_v^+(T) \leq (1+\epsilon)A(T, v, \epsilon).$$

The running time of the algorithm is polynomial in n/ϵ (hence of the form $O(n^a/\epsilon^b)$ for some fixed constants $a > 0$ and $b > 0$).

Our proof is constructive in the sense that we actually describe the algorithm. We remark (see Section 3.1) that the algorithm extends almost without change to estimating the cover and return time of arbitrary Markov chains on trees, though the running time in this case is polynomial in the cover and return time itself rather than in the number of states. (This distinction was not necessary for simple random walks on trees because there the cover time is bounded by a polynomial in the number of vertices.) The algorithm also extends to the case when we are given a set S of vertices in the tree, and are required to estimate the expected time by which a random walk on T covers the vertices of S and returns to v. See Section 3.2.

The additive error in the approximation provided by Theorem 1.1 is at most $\epsilon C_v^+(T)$. As $C_v^+(T) < 2n^2$ for every n-vertex tree, see Section 1.1, it follows that the additive error is at most $\epsilon/2n^2$. The running time of the algorithm remains polynomial in n even if $\epsilon < 1/2n^2$, and hence Theorem 1.1 also provides approximations of the cover and return time with arbitrarily small additive error.

The proof of Theorem 1.1 as appears in Section 2 applies to the cover and return time but not to the cover time. It is possible to use the cover and return time in conjunction with the hitting times from leaves of T to v in order to obtain accurate estimates on the cover time. Further, the algorithm in Theorem 1.1 and its proof can be adapted to handle also the cover time. Hence a statement similar to that of Theorem 1.1 (see Theorem 3.4) applies also to the cover time. We sketch the proof in Section 3.3.

2 The deterministic algorithm

Many computational problems that are difficult to solve on graphs are easy (polynomial time solvable) on trees. The algorithmic paradigm that is often used in these cases is dynamic programming. We shall also use dynamic programming so as to approximate the cover time on trees. The difficulty is that the cover time per se is not a quantity that lends itself well to aggregation of information. For example, consider a tree T with root vertex r connected to two vertices r_1 and r_2, which are root vertices of subtrees T_1 and T_2. Even if one is given the complete distribution function for the cover and return time of the subtrees T_1 and T_2, it is not immediately clear (to the authors) how to combine this information so as to obtain $C_r^+(T)$. To overcome this difficulty, we extend an approach that was used by Aldous [1] for evaluating the cover time of balanced trees.

Let T be an arbitrary tree with vertices $v_1, \dots v_n$ on which we wish to estimate $C_1^+(T)$. For the sake of uniformity of the notation, we shall introduce a new root vertex r to the tree connected only to v_1, thus obtaining a new tree that we shall call T_r. The tree T_r has $n+1$ vertices and n edges. For the sake of establishing notation, orient all edges away from the root, and for every $1 \leq i \leq n$, let e_i be the unique edge whose endpoint is v_i. As a convention, we say that e_i is *traversed* whenever the walk enters v_i through e_i (but not when the walk exits v_i through e_i). Let T_i be the subtree rooted at v_i (hence $T_1 = T$). Now we define the key quantity on which we shall employ dynamic programming.

Definition 2.1 *Using the notation introduced above, for $1 \leq i \leq n$ and for $t \geq 1$, let $P_i(t)$ denote the probability that a walk on T_r that starts at v_i visits all vertices of T_i before edge e_i is traversed t times.*

As a simple example, if v_i is a leaf of T, then $P_i(t) = 1$ for every t. This will serve as the base case that will start off our dynamic programming. Our goal will be to compute $P_1(t)$ for all t. Using these values, we may consider $E(1) = \sum_t (1 - P_1(t))$ which is equal to the expected number of times that e_1 is traversed in a walk on T_r that starts at v_1 and covers T_1. (Technically, $E(1)$ is an infinite sum. However, the sum converges since necessarily $E(1) < C_1^+(T_r)$, and $C_1^+(T_r) \leq 2n^2$.)

The following lemma shows the connection between the value of $E(1)$ and the desired $C_1^+(T)$.

Lemma 2.2 *With notation as above, $C_1^+(T) = 2(n-1)E(1)$.*

Proof: Let $C[v_1, r]$ denote the *commute time* between v_1 and r in T_r (the expected number of steps it takes a walk that starts at v_1 to visit r and return to v_1). As mentioned in Section 1.1, $C[v_1, r] = 2n$. Observe that due to Wald's lemma, $C_1^+(T_1) = E(1)C[v_1, r]$. (An intuitive way to see the latter equality is by considering an extremely long random walk on T_r that starts at v_1, and covers T_1 many times. Break the walk into segments that correspond to the walk covering T_1 and returning to v_1. During the first ℓ such segments, with ℓ large, the ergodic theorem implies that the number of commutes to r is close to $E(1)\ell$. Taking $\ell \to \infty$ then yields the identity.) Moreover, observe that one can relate the cover time of T_1 in T_r to that in T by subtracting the steps along the edge e_1 (in both directions). Linearity of expectation then implies that $C_1^+(T) = C_1^+(T_1) - 2E(1)$. Putting everything together we deduce that $C_1^+(T) = 2(n-1)E(1)$. $\square$

As noted above, to compute $E(1)$ from $P_1(t)$ involves an infinite sum. To obtain a finite algorithm, we shall truncate the sum when t exceeds a sufficiently large value N. To keep the presentation simple, we shall not attempt to optimize the value of N here (not for trees in general and not for any tree specifically), but just note that N can be chosen to be $O(n^2 \log 1/\epsilon)$, because it is not hard to show that for some universal constant $c > 0$,

$$P_1(t) \geq 1 - e^{-ct/n^2} . \tag{1}$$

Indeed, since $C_1^+(T_r) \leq 2n^2$, the probability to cover T_r within the first $4n^2$ traverses of the edge (r, v_1) is at least $1/2$, which implies the estimate on $P_1(t)$. More generally, taking N as $O(n^2 \log(n + \frac{1}{\epsilon}))$ we will be able for every vertex v_i to consider the values of $P_i(t)$ only for t up to N, while still eventually achieving a $(1 \pm \epsilon)$ multiplicative approximation for $C_1^+(T)$.

We now proceed to describe an exact dynamic programming procedure in an idealized world in which computations can be done with arbitrary precision and summations may include infinitely many summands (though all sums do converge). Later we shall discuss how the dynamic programming can be carried out in polynomial time with only a small loss in the accuracy of the computations.

For every vertex v_i we shall compute the infinite vector $P_i = \{P_i(t)\}$ for all values of t. (Needless to say, in our actual algorithm we shall truncate this vector an $t = N$.) As noted, for every v_i that is a leaf of T_r, this is the all 1 vector. For every other vertex v_i, let D_i denote the set of direct descendants of v_i (those vertices connected to v_i by edges other than e_i). Given that T_r is a tree, it will always be the case that if we have not yet computed P_1, then there is some vertex v_i for which P_i has not yet been computed but the vectors P_j were already computed for all $v_j \in D_i$. Hence we will compute P_i for such a vertex v_i and make progress. The computation will involve quantities that shall be defined next.

Fix a vertex v_i of interest. To simplify notation, let $d = |D_i|$ be the number of direct descendants of v_i. Rename them as $u_1, \dots, u_d$.

Definition 2.3 *For vertex v_i, $t \geq 1$, $t_1 \geq 1, \dots, t_d \geq 1$, define $Q_i(t_1, \dots, t_d; t)$ to be the probability that in a walk on T_r that starts at v_i, each edge (v_i, u_j) is* *traversed exactly t_j times (in the direction into u_j) before the edge e_i is traversed t times.*

Observe (though we shall not need to use this fact) that for two vertices v_i and v_j with the same number of descendants, the functions Q_i and Q_j are identical.

We now have a recursive formula for $P_i(t)$ in terms of the vectors P_j for the descendants $v_j \in D_i$. (So as to keep notation simple, we use in this formula the convention that P_i refers to v_i, but P_j refers to u_j rather than v_j.)

$$P_i(t) = \sum_{t_1 \geq 1, \ldots, t_d \geq 1} Q_i(t_1, \ldots, t_d; t) \prod_{j=1}^{d} P_j(t_j) \tag{2}$$

Let us explain Equation (2). We wish to compute the probability that T_i is covered before e_i is traversed t times. In order to cover T_i, each vertex of $u_j \in D_i$ must be visited at least once, and the subtree T_j rooted at u_j needs to be covered. Once we fix the stopping condition of the edge e_i being traversed t times, the distribution of the number of visits (from their parents) to the descendants $u_j \in D_i$ is given by the function Q_i. Subtree T_j needs to be covered by the time the edge (v_i, u_j) is traversed t_j times, one of which is the first entry to u_j, and hence the term $P_j(t_j)$ gives the probability of T_j being covered. We can take the product of the terms $P_j(t_j)$, because the walks within different subtrees are independent.

Using Equation 2 and the fact that the vectors P_i are known for all leaves, we get an inductive definition for P_1, and then Lemma 2.2 can be used to compute $C_1^+(T)$. However, there are several obstacles to obtaining a polynomial time algorithm. We list these obstacles, and then explain how to overcome them, paying only a multiplicative factor of $(1 \pm \epsilon)$ in the accuracy of the computation.

1. **Range of summation.** Each variable t_j ranges over infinitely many values. As explained earlier, this will be handled by limiting the range between 1 and N for sufficiently large N.

2. **Combinatorial explosion.** Even if the range of the summation of each variable is limited to N, the number of terms in the summation is N^d. Since d need not be bounded by a constant (the tree may have vertices of arbitrarily large degrees), this number will not be polynomial in n. We shall refine the dynamic programming approach so as to overcome this obstacle.

3. **Finite precision.** Computation cannot be performed with infinite precision. We shall either need to show that the numbers involved can always be represented using polynomially many bits, or round some of the numbers and account for the error introduced by the rounding.

It would be more convenient for us to first deal with the second obstacle, and only later with the other obstacles.

2.1 Avoiding the combinatorial explosion

For every vertex v_i, if $|D_i| > 2$, construct an arbitrary binary tree B_i (each internal node has two children) with $d = |D_i|$ leaves, placing v_i at the root and $u_1, \ldots, u_d$ at the leaves. There are $d-2$ internal nodes (in addition to the root) that we shall name as $b^i_1, \ldots, b^i_{d-2}$. For simplicity of notation, let us fix the structure of the tree to be a path $v_i, b^i_1, \ldots b^i_d$, with u_1 connected to v_i, u_d connected to b^i_{d-2}, and u_j for $1 < j < d$ connected to b^i_{j-1}.

The random walk on T_r can be simulated as follows. Whenever the walk on T_r reaches v_i, with probability $1/(d+1)$ it takes the edge e_i, and with probability $d/(d+1)$ it goes to one of the children, chosen uniformly at random. This random choice of child is simulated by a walk on the tree B_i. Conditioned on having decided not to take the edge e_i, at every internal node of the tree B_i (including the root), choose one of the two children with probability proportional to the number of leaves of B_i that are descendants of the child. For example, at internal node b^i_k with $k < d-2$, go to leaf v_{k+1} with probability $1/(d-k)$ and to internal node b^i_{k+1} with probability $(d-k-1)/(d-k)$. It can readily be seen that each leaf is reached with the same probability. Being at a leaf u_j in T_r and deciding to take the edge (u_j, v_i) is simulated in T_B by taking the path u_j to v_i in the tree B_i.

For the simulated random walk, every vertex has only two children. This is the key to avoiding the combinatorial explosion. Observe that building such trees B_i for all vertices v_i, we change T_r into a tree T_B which is a subtree of the binary tree. Every leaf of T_B is a leaf of T_r, and so it follows that the total number of vertices in T_B is at most $2n$.

The tree T_B is still rooted at r like T_r, and r has degree 1 also in T_B. Except for r, T_B has two types of vertices: those which were original vertices of T (and were denoted by v_i), and those that were added by the subtrees B_i (and were denoted by b^i_k). For uniformity of notation, we use w_i to denote vertices of T_B, regardless of the origin of the vertex. However, we associate with each vertex w_i a weight W_i. The weight of each of the original vertices of T_r is 1. The weight of a vertex b^i_k of B_i is always greater than 1, and equal to the number of leaves of B_i in the subtree of B_i rooted at b^i_k. (With the notation that we used above, it turns out that this weight is equal to $d-k$.) As in the case of T_r, we now use T_i to denote the subtree of T_B rooted at w_i.

Recall that a walk on a graph is a sequence of vertices (that respects the adjacency structure of the graph). We now define a (random) walk $\{S_n\}$ on T_B, as follows.

1. At r, move to its unique neighbor w_1.
2. Let the walk be at a vertex w_i with $W_i = 1$ (hence, an original vertex of T). Let w_p be its parent node.
    (a) If w_i is a leaf, move to its parent vertex w_p.
    (b) If w_i has only one child, move to this child with probability $1/2$ and to w_p with probability $1/2$.

(c) Otherwise, w_i must have exactly two children, one of them (say w_l) of weight 1 and the other (say w_r) of weight $W_r \geq 1$. Move to w_p with probability $1/(2+W_r)$, to w_l with probability $1/(2+W_r)$, and to w_r with probability $W_r/(2+W_r)$.

3. Let the walk be at a vertex w_i with $W_i > 1$ (hence a vertex that was introduced through some subtree B). Let w_p be its parent node, and w_l and w_r be its two children. At least one of these children is an original vertex of T_r, hence we assume without loss of generality that $W_l = 1$.

   (a) If w_i was last entered from one of its children, move to w_p.

   (b) If w_i was last entered from w_p, move to w_l with probability $1/(1+W_r)$ and to w_r with probability $W_r/(1+W_r)$.

Note that the random walk thus defined is not Markovian, while the process $\{(S_{n-1}, S_n)\}_{n\geq 1}$ is Markovian.

So far, we have defined two random walk processes, one on T_r and one on T_B. For a walk on T_B, we now define the *projection* of the walk to be the subsequence of vertices that includes only the original vertices of T_r (removing the vertices introduced by the subtrees B from the sequence). Random walks on T_B simulate random walks on T_r in the sense that the projection of a random walk on T_B is precisely a random walk on T_r.

Definition 2.1 applies with minor changes to walks on T_B. We present the revised definition.

Definition 2.4 *For vertex w_i in T_B, with parent vertex denoted by w_p, and for $t \geq 1$, define $P_i(t)$ to be the probability of the following event:*

1. *If $W_i = 1$, the event is that a walk on T_B that starts at w_i visits all vertices of T_i before traversing the edge (w_p, w_i) t times.*

2. *If $W_i > 1$, the event is that a walk on T_B that just entered w_i from w_p visits all vertices of T_i before traversing the edge (w_p, w_i) t additional times.*

Likewise, Definition 2.3 needs to be modified so as to account for the existence of different types of vertices in T_B.

Definition 2.5 *For vertex $w_i \neq r$ in T_B, let w_p denote its parent vertex and let w_l and w_r denote its two children (or only w_l if w_i has one child). For $t \geq 1$, $t_l \geq 1$, $t_r \geq 1$, define $Q_i(t_l, t_r; t)$ (or $Q_i(t_l; t)$ if w_i has only one child) to be the probability of the following event:*

1. *If $W_i = 1$, then the event is that in a walk on T_B that starts at w_i, edge (w_i, w_l) is traversed exactly t_l times and edge (w_i, w_r) is traversed exactly t_r times before the edge (w_p, w_i) is traversed t times. (If w_i has only one child, then remove the condition on t_r.)*

2. *If $W_i > 1$, then the event is that in a walk on T_B that just entered w_i from w_p, edge (w_i, w_l) is traversed exactly t_l times and edge (w_i, w_r) is traversed exactly t_r times before the edge (w_p, w_i) is traversed t additional times.*

Armed with the new definitions for P_i and Q_i, Equation (2) when applied to T_B simplifies to:

$$P_i(t) = \sum_{t_1 \geq 1, t_2 \geq 1} Q_i(t_1, t_d; t) P_1(t_1) \cdot P_2(t_2) \tag{3}$$

Inductively applying Equation (3) in T_B we obtain the vector P_1 in T_B, which is equal to the vector P_1 in T_r.

2.2 Limiting the range of summation

For some sufficiently large value of N (to be determined later), we shall truncate all vectors P_i after N entries, implicitly assuming that $P_i(t) = 1$ for all $i \geq N$. Hence we shall set $P_i(N) = 1$, regardless of its true value or computed value. For vertices w_i that are not leaves, this certainly introduces an error. Moreover, this error propagates and amplifies through our use of Equation (3).

We shall modify Definition 2.5 to reflect the fact that we no longer distinguish between different values of t_j that are larger than N.

Definition 2.6 *For vertex $w_i \neq r$ in T_B, let w_p denote its parent vertex and let w_l and w_r denote its two children (or only w_l if w_i has one child). For $1 \leq t \leq N$, $1 \leq t_l \leq N$, $1 \leq t_r \leq N$, define $Q_i(t_l, t_r; t)$ (or $Q_i(t_l; t)$ if w_i has only one child) to be the probability of the following event:*

1. *If $W_i = 1$, then the event is that in a walk on T_B that starts at w_i, edge (w_i, w_l) is traversed exactly t_l times (and at least t_l times in the special case that $t_l = N$) and edge (w_i, w_r) is traversed exactly t_r times (and at least t_r times in the special case that $t_r = N$) before the edge (w_p, w_i) is traversed t times. (If w_i has only one child, then remove the condition on t_r.)*

2. *If $W_i > 1$, then the event is that in a walk on T_B that just entered w_i from w_p, edge (w_i, w_l) is traversed exactly t_l times (and at least t_l times in the special case that $t_l = N$) and edge (w_i, w_r) is traversed exactly t_r times (and at least t_r times in the special case that $t_r = N$) before the edge (w_p, w_i) is traversed t additional times.*

We can now modify our recursive formula to have only finitely many terms. It no longer computes the true value of $P_i(t)$, so we shall call the quantity that it computes $P_i^1(t)$. The function Q to be used in this formula is the one from Definition 2.6. $P_i^1(N)$ is not computed by this formula, but instead set to 1.

$$P_i^1(t) = \sum_{1 \leq t_l \leq N, 1 \leq t_r \leq N} Q_i(t_l, t_r; t) P_l^1(t_l) P_r^1(t_r) \tag{4}$$

This completes the description of how we limit the range of summation to be finite. We now analyze the effect of this approximation. For a given choice of N, let $\delta > 0$ be such that for every i, $(1+\delta)P_i(N) \geq 1$. For concreteness, take

$$\delta = 2(1 - P_i(N)) \leq 2e^{-cN/n^2}\,, \tag{5}$$

see (1). We shall express the relative error in the approximation as a function of N and δ, and thereafter choose N such that together with the implied δ, the relative error is smaller than ϵ.

At the leaves of T_B there is no error in the respective vector P_i. At a vertex w_i whose two children (or single child, if w_i has only one child) are leaves, a multiplicative error of at most $(1+\delta)$ is introduced because $P_i^1(N)$ is rounded to 1, even though its true value may have been $1/(1+\delta)$. Consider now some other arbitrary vertex w_i, let w_l and w_r be its children, and let $(1+\delta_l)$ and $(1+\delta_r)$ be upper bounds on the multiplicative errors in any of the entries of the vectors P_l^1 and P_r^1. Then by inspection of Equation (4), the multiplicative error in any entry of P_i^1 is at most $(1+\delta_l)(1+\delta_r)$. Since T_B has at most $2n$ vertices, it follows that the multiplicative error at entries of P_1^1 (compared to the true entries of P_1) is at most $(1+\delta)^{2n}$.

Recall that we needed the vector P_1 so as to compute the expectation $E(1) = \sum_{t\geq 1}(1 - P_1(t))$. Instead we now compute at approximation $E^1(1) = \sum_{1\leq t\leq N}(1 - P_1^1(t))$. Hence our total error in this computation is:

$$E(1) - E^1(1) = \sum_{1\leq t\leq N}(P_1^1(t) - P_1(t)) + \sum_{t>N}(1 - P_1(t))$$

The first of these summations is at most $N((1+\delta)^{2n} - 1)$. If $\delta \ll \frac{1}{2n}$ then this value is approximated well by $2nN\delta$. In the second of these summations, each term is of value at most δ. Moreover, for every t, $(1 - P_1(t+N)) \leq \delta(1 - P_1(t))$. Hence if $\delta < 1/2$ then the second summation can be upper bounded by a geometric series of sum $2N\delta$. Hence the total additive error is at most $2(n+1)N\delta$, and we wish it to be smaller than $\epsilon E(1)$. It is not hard to show that in every tree $E(1) \geq 1$ (in fact, in every tree $E(1)$ is essentially the cover time divided by $2n$, and the cover time of a graph is $\Omega(n \log n)$), and hence we shall simplify the desired inequality to $nN\delta \leq \epsilon$. This requires choosing N such that $\delta \leq \frac{\epsilon}{nN}$. With the choice of δ in (5), a value of $N = cn^2 \log(n + \frac{1}{\epsilon})$ for a sufficiently large constant c would suffice for all trees.

We remark that in this paper we just give a sufficient value of N. Much lower values of N will also work for special families of trees (essentially, a factor of n can be replaced by their cover time divided by n), and moreover, we need not use the same value of N for all vertices of T_B (in particular, for the leaves we may take $N = 1$). These kind of optimizations are omitted from this paper.

2.3 Computation with finite precision

Having established the value of N for which P_1^1 is a sufficiently close approximation for P_1, it remains to verify that P_1^1 can indeed be computed in polynomial time. For this, we need to be able to compute the values $Q_i(t_l, t_r; t)$. Let us first observe that the role of vertex w_i in the value of this expression is only in determining the weight of w_r (the weight of w_l is always 1). Definition 2.6 offers several cases for the definition of Q_i, and we shall address only some of them here. The other cases are handled similarly.

Let us compute $Q_i(t_l, t_r; t)$ when $W_i = 1$, $t_l \neq N$ and $t_r \neq N$. Whenever the walk is at w_i, it has probability $p_p = 1/(2 + W_r)$ to go to w_p, probability $p_l = 1/(2 + W_r)$ to go to w_l, and probability $p_r = W_r/(2 + W_r)$ to go to w_r. The probability of exactly t_l visits to w_l and exactly t_r visits to w_r prior to t visits to w_p is exactly

$$\binom{t + t_l + t_r - 1}{t_l}\binom{t + t_r - 1}{t_r}(p_p)^t (p_l)^{t_l} (p_r)^{t_r} .$$

The upper bound of N implies that both the numerator of this expression and the denominator are numbers that can be expressed by $O(N(\log N + \log n))$ bits.

If $t_r = N$ then $Q_i(t_l, N; t)$ will be computed differently. First, ignoring moves into w_r (as if w_i has only one child), compute $Q_i(t_l; t)$. Then subtract $\sum_{0 \le t_r \le N-1} Q_i(t_l, t_r; t)$ to get the desired result. Observe that for the final answer one can use a common denominator $(N!)^2(2 + W_r)^N$, and hence still expressible in a polynomial number of bits.

A similar argument applies to the computation of $Q_i(t_l, t_r; t)$ when $W_i > 1$, $t_l \neq N$ and $t_r \neq N$. In this case, the probability of exactly t_l visits to w_l and exactly t_r visits to w_r prior to t additional visits to w_p vanishes unless $t = t_l + t_r$, in which case it equals

$$\binom{t}{t_l}\frac{W_r^{t_r}}{(1 + W_r)^t} ,$$

where W_r is the weight of the right descendent of W_i.

Following the computation in Equation (4), and thereafter applying it to all vertices of T_B, one sees that one can obtain a rational number with denominator $(N!n!)^{O(n)}$, and likewise with a numerator expressible by polynomially many bits. Hence in principle, all computations can be performed exactly in polynomial time, though they would be very tedious.

A more practical approach is to round the numbers to numbers of shorter representations. Clearly, this can be done while maintaining the relative error in the range $(1 \pm \epsilon)$, but we omit concrete suggestions of how to do this.

3 Extensions

We present in this section several extensions of Theorem 1.1.

3.1 Arbitrary Markov chains on trees

Consider a Markov chain $\{S_t\}$ with state space the vertices of a (finite) tree T, where transitions are allowed only between neighbors. Because of the tree structure, the Markov chain is reversible, and hence there exist conductances $\mathcal{C}_{\{u,v\}}$ (with u, v neighboring vertices in the tree) such that the transition probability from u to v equals $\mathcal{C}_{\{u,v\}} / \sum_{w:w\sim u} \mathcal{C}_{\{u,w\}}$. Let $\mathcal{R}_{\{u,v\}} = 1/\mathcal{C}_{\{u,v\}}$ denote the respective resistance between neighboring vertices in the tree, and assume first that all resistances are integer valued. Consider the tree T' in which each edge $\{u, v\}$ is replaced by a chain of length $\mathcal{R}_{\{u,v\}}$, and a simple random walk $\{RW_t\}$ on T'. Thus, each vertex of T corresponds to a vertex of T'. Further, the random walk $\{RW_t\}$ induces a Markov process on T, and the transition probabilities of the latter coincide with those of $\{S_t\}$. In particular, the quantity $E(1)$ corresponding to T is identical to that corresponding to T', and can be computed accurately by Theorem 1.1. For the tree T, we have $C[v_1, r] = 2\sum \mathcal{R}_{\{u,v\}}$, and $C_1^+(T) = (C[v_1, r] - 2)E(1)$ by an adaptation of Lemma 2.2. We conclude from these facts that the cover and return time can be computed by the algorithm of Theorem 1.1, with the running time polynomial in the cover and return time itself rather than in the number of states. It is straightforward to approximate the above in case the resistances are not integer-valued.

3.2 Covering a specified set of vertices

Let T' be a subtree of T, rooted at v_1. The algorithm of Theorem 1.1 applies equally well to the evaluation of the cover and return time of T' by a random walk on T, denoted $C_1^+(T'; T)$, as follows. The quantity $E(1)$ for T', denoted $E_{T'}(1)$, can be computed by the algorithm (applied to T'). We then have (again, by an adaptation of Lemma 2.2) that $C_1^+(T'; T) = 2(n - 1)E_{T'}(1)$.

3.3 Computing the cover time

Fix a tree T and a starting vertex v. Given a vertex u, let $P_{last}[u]$ denote the probability that for a random walk on T that starts at v, the last vertex to be visited is u. Clearly, $P_{last}[u] \neq 0$ iff u is a leaf of T (different than v). Let $H[u, v]$ denote the expected hitting time in T from u to v. Then the cover time satisfies:

$$C_v(T) = C_v^+(T) - \sum_u P_{last}[u]H[u, v]$$

Recall that for every vertex u, $H[u, v]$ can be computed exactly in polynomial time (moreover, the known algorithms compute $H[u, v]$ for all u simultaneously, though this fact is not needed here), and that $C_v^+(T)$ can be computed with arbitrary small additive error. It follows that it suffices to estimate the quantities $P_{last}[u]$ with sufficiently high precision in order to obtain an accurate estimate of the cover (without return) time.

The latter task can be performed in a way similar to that described in Theorem 1.1. We sketch the steps, assuming a reduction to a binary tree T_B has already been performed as in Section 2.

We begin with a definition.

Definition 3.1 *For a vertex v_i with $u \in T_i$ and for $t \geq 1$, let $A_i(t)$ denote the probability that a walk on T_B that starts at v_i satisfies the following conditions.*

- *It does not visit u before edge e_i is traversed $t-1$ times.*
- *It does visit u by the time edge e_i is traversed t times.*
- *u is the last vertex from T_i to be visited.*

Clearly, $P_{last}[u] = \sum_t A_v(t)$, where v is the starting vertex of the walk and $u \in T_v$.

The following definition is similar to Definition 2.1. It will be used later in situations where $u \notin T_i$.

Definition 3.2 *For a vertex v_i and for $t \geq 1$, let $P_i(t)$ denote the probability that a walk on T_B that starts at v_i visits all vertices of T_i before edge e_i is traversed t times.*

Note that we have already seen in Section 2 how all $P_i(t)$ can be computed. We now explain how this can be used in order to compute all $A_i(t)$.

Definition 3.3 *For a vertex v_i with two children (v_l and v_r), with $u \in T_{v_l}$, $t \geq 1$, $t_l \geq 1$, $t_r \geq 1$, let $R_i(t_l, t_r; t)$ denote the probability that a walk on T_B that starts at v_i satisfies the following conditions.*

- *By the time edge e_i is traversed $t-1$ times, the edge e_l is traversed at most $t_l - 1$ times.*
- *By the time edge e_i is traversed t times, the edge e_l is traversed at least t_l times.*
- *By the time edge e_l is traversed t_l times, the edge e_r is traversed exactly t_r times.*

The function $R_i(t_l, t_r; t)$ can be computed efficiently in a way similar to that described in Section 2.3. The details are tedious and are omitted. Now, with u as in Definition 3.3, $A_i(t)$ can be computed using the following recursive formula:

$$A_i(t) = \sum_{t_l \geq 1, t_r \geq 1} R_i(t_l, t_r; t) A_l(t_l) P_r(t_r)$$

The truncation of the sum to a finite sum can performed as in Section 2.2, with a similar computational cost.

The outline above (together with additional technical details which are omitted) implies the following theorem.

Theorem 3.4 *There is a deterministic algorithm that given a tree T on n vertices, a starting vertex v and a slackness parameter $\epsilon > 0$, outputs a value $A(T, v, \epsilon)$ that approximates the cover time $C_v(T)$ within a factor of $1 \pm \epsilon$. Namely,*

$$(1-\epsilon)A(T, v, \epsilon) \leq C_v(T) \leq (1+\epsilon)A(T, v, \epsilon).$$

The running time of the algorithm is polynomial in n/ϵ.

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